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Dynamic Race Detection in Parallel Programs

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Abstract

Recent hardware developments have pushed parallel computing out of the niche of numeric applications into the mainstream. Unfortunately, parallel programs may contain synchronization defects, a class of defect which is difficult to detect. A significant source of these defects is the phenomenon of *data races*, i.e., unsynchronized accesses to shared data. Since parallel programs are schedule-dependent, reproducing data races is often difficult. Programs encountering data races often do not crash immediately, resulting in mysterious and unpredictable behavior.

Currently, available tools tend to miss many data races, or to produce an overwhelming number of false alarms, regardless of whether they are based on static analysis or dynamic analysis. Both types of analysis also have their own specific problems. Static analysis, due to the state explosion problem, is not applicable to large programs, or, alternatively, the analysis has to focus on a small subset of fault types. Dynamic analysis, on the other hand, is limited to finding faults in the code which is actually executed. Additionally, dynamic analysis is either very slow or reports numerous false warnings.

In this work, we propose a dynamic approach for race detection based on a synthesis of *lockset* and *happens-before* analyses. The approach provides a lower rate of both false positives and false negatives (missed races). The basic idea is to consult the happens-before relation whenever the lockset algorithm indicates a possible race. The increased precision is due to more detailed state machines and adjustment of the sensitivity of the detector for different kinds of applications. Additionally, a technique to correctly handle *inter-thread event notifications* further improves the accuracy of the detector.

Furthermore, we present a new method to deal with *ad-hoc synchronizations*, i.e., programmer-defined synchronizations in source code. The method is also able to identify synchronization operations from unknown libraries, resulting in a *universal race detector.*

Our race detection approach is automatic, without any user intervention or reliance on source code annotation, and has been implemented as a tool, which we named *Helgrind*$^+$. Results from several benchmarks demonstrate a significant reduction in false positive rates and false negative rates compared to existing race detectors, with a negligible increase in overhead.

Kurzfassung

Parallele Programme sind anfällig für Synchronisierungsfehler, die schwierig aufzuspüren sind. Ursache sind meist sog. Wettlaufsituationen, d.h. unsynchronisierte Zugriffe auf gemeinsam genutzte Daten. Solche Wettlaufsituationen lassen sich nicht zuverlässig reproduzieren, da ihr Auftreten in der Regel abhängig von einer konkreten Ausführungsreihenfolge ist. Erschwerend kommt hinzu, dass sie sich oft nicht in unmittelbaren Programmfehlern manifestieren, sondern zu unvorhersehbarem Programmverhalten führen, dessen Ursache dann nur noch schwer zurückzuverfolgen ist.

Verfügbare Wettlauferkenner, sowohl statische als auch dynamische, neigen dazu, Wettläufe zu übersehen, oder eine Flut von Falschmeldungen zu liefern. Hinzu kommen Probleme, die in der Art der Analyse selbst begründet sind: Die statische Analyse kann häufig nicht auf größere Probleme angewendet werden, da die Anzahl der zu untersuchenden Programmzustände exponentiell wächst. Dynamische Analyse hingegen kann nur Fehler in solchen Programmteilen finden, die tatsächlich ausgeführt werden. Auch wird durch dynamische Analyse die Programmausführung erheblich verlangsamt.

Mit *Helgrind*$^{+}$ stellen wir einen neuen, dynamischen Ansatz vor, der Lockset- und Geschieht-Vorher-Analyse kombiniert: Die Geschieht-Vorher-Beziehung wird immer dann herangezogen, wenn der Lockset-Algorithmus einen Wettlauf meldet. Durch Anpassen der internen Zustandsautomaten – auch unter Berücksichtigung verschiedener Programmklassen – können wir Genauigkeit und Qualität der Wettlauferkennung deutlich erhöhen. Dazu trägt auch eine neue Technik bei, die den Signalaustausch zwischen parallelen Fäden korrekt handhabt.

Nicht zuletzt gelingt es uns, benutzerdefinierte Synchronisierung, sogenannte adhoc-Synchronisierung, zu berücksichtigen. Da wir so auch Synchronisierungsprimitive von nicht direkt unterstützen Bibliotheken zu erkennen vermögen, stellt unser Ansatz den ersten *universellen Wettlauferkenner* dar.

Für die Analyse sind keine weiteren Benutzereingaben und keine Annotationen des Quellcodes erforderlich. Anhand von Benchmarks können wir im Vergleich zu existierenden Wettlauferkennern eine deutliche Reduzierung sowohl der Zahl der nicht gemeldeten Wettläufe als auch der Zahl der Falschmeldungen nachweisen. Dabei bleibt der zusätzliche Mehraufwand bei Speicherverbrauch und Ausführungszeit vernachlässigbar.

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Chapter 1.

Introduction

1.1. Motivation

In the past years, software was able to benefit from the increased CPU clock rate. Nowadays, because of some limitations such as heat dissipation problems, processing power is increased by integrating multiple processors into one chip, introducing multi-core and many-core chips. Programs cannot benefit from additional cores unless the program tasks are executed in parallel to utilize them. But writing correct parallel programs is a difficult task. Due to the non-deterministic behavior of parallel programs, new types of defects occur that are not easy to detect. More prevalent among them are *data races* that lead to inconsistent data. A *data race* occurs in a multi-threaded program when two threads access the same memory location with no ordering constraints enforced between the accesses, and at least one of the accesses is a write [34].

The following example illustrates a data race that may result in inconsistent data because of non-deterministic orderings of parallel threads.

```
TEMP1 = X
TEMP1 = TEMP1 + 1
X = TEMP1
```

(a) Thread 1

```
TEMP2 = X
TEMP2 = TEMP2 + 1
X = TEMP2
```

(b) Thread 2

Figure 1.1.: Data race example.

In Figure 1.1, two concurrent accesses to a shared location `X` cause an undesired data race. Each thread increments the shared variable `X` by using its thread local variable `TEMP`. The variable `X` is initialized to zero at the beginning. If both threads run in parallel, the result assigned to variable `X` depends on the actual execution order of operations, i.e. the thread interleaving. As an example, consider the following two possible execution orders:

First Interleaving: In the first interleaving, Thread 1 initializes the value of TEMP to X and then increments it. At the same time, Thread 2 reads the value of X and assigns it to its other local variable TEMP and increments it. In this case, X becomes one at the end (X==1).

```
TEMP1 = X
TEMP1 = TEMP1 + 1


X = TEMP1
```

(a) Thread 1

```

TEMP2 = X
TEMP2 = TEMP2 + 1

X = TEMP2
```

(b) Thread 2

Figure 1.2.: First interleaving increments X once.

Second Interleaving: In the second interleaving, after Thread 1 increments the value of TEMP and assigns it to X, Thread 2 reads X and increments it: We get a different result (X==2).

```
TEMP1 = X
TEMP1 = TEMP1 + 1
X = TEMP1
   .
   .
   .
```

(a) Thread 1

```
   .
   .
   .
TEMP2 = X
TEMP2 = TEMP2 + 1
X = TEMP2
```

(b) Thread 2

Figure 1.3.: Second interleaving increments X twice.

The behavior of this program is non-deterministic, because we get different outputs (X==1 or X==2). If we want to change the program, so that it results in consistent output, we have to make the increment operation on the variable X atomic. That is, no other thread should be allowed to change the value of X, while Thread 1 is executing and working on X.

A possible solution for this simple program is to use critical sections or locks. Only one thread in Figure 1.4 can enter the critical section between *lock(l)* and *unlock(l)* at any point of time. Hence, we get rid of the undesired results and prevent a data race. Generally, in order to prevent undesired concurrent accesses to shared locations, we must explicitly synchronize threads by the means of synchronization primitives such as *lock/unlock* or *signal/wait.*

Although locking policies are commonly used for synchronization in concurrent programs, it is also easy to get it wrong. The user may miss to lock critical

```
lock(l)
  TEMP = X
  TEMP = TEMP + 1
  X = TEMP
unlock(l)
```

(a) Thread 1

```
lock(l)
  TEMP = X
  TEMP = TEMP + 1
  X = TEMP
unlock(l)
```

(b) Thread 2

Figure 1.4.: Elimination of data race by lock protection.

parts of the program. Another common problem arises from circular dependencies between locks that may result in deadlocks. In this case, we need tools to detect the fault and produce a proper warning. On the other hand, the programmer may overdo using synchronization primitives to achieve correctness which makes programs inefficient. So, we may also need tools to remove excessive synchronizations.

Finding Synchronization Defects

Different methods have been proposed to find synchronization defects in parallel programs. They can be classified into two categories: static methods and dynamic methods. Static methods that check the program code and employ compile-time analysis of the program source do not scale with program size. They are still infeasible for programs larger than a few 10,000 lines of code. The approach produces excessive false positives, since compile-time analysis cannot understand the semantics of the program, and is unable to determine the precise set of possible thread interleavings. Thus, it makes conservative estimates.

On the other hand, dynamic methods are based on runtime checking and applicable to programs of any size. But they are only able to find synchronization defects in parts of the program that are actually executed. To compensate for this drawback, it is necessary to have sufficient program coverage during the test runs in order to find possible faults. Furthermore, tracing mechanisms slow down programs. It is possible to reduce the overhead by logging all necessary events and performing the analysis offline.

Dynamic methods for data race detection fall into two main categories: *lockset* and *happens-before*. Lockset methods check if a shared variable is protected correctly by locks and the program obeys a locking discipline. Happens-before methods verify whether accesses to a shared variable are ordered by synchronization notifications between threads. Both methods have some shortcomings: The lockset algorithm is scalable and insensitive to interleaving but produces lots of false positives. This is because it considers only the lock primitives. Comparatively, happens-before methods are sensitive to interleavings and not scalable but produce fewer false positives.

A practical dynamic race detector must have good scalability and performance, combined with a low false alarm rate. It should not be sensitive to thread interleaving. Our goal in this work is to provide a practical race detector that combines the two methods so as to get the strengths of both. We present a new race detection approach that detects potential races, not just races that actually occur in executions. Since modern parallel software is getting extremely complex, a practical race detector is important and quite useful.

1.2. Problem Statement

A dynamic race detector detects synchronizations based on explicit calls of synchronization primitives in a program. The race detector examines if synchronizations are consistently used and reports any inconsistent use of synchronization or missing synchronization as a data race. Race detection is done automatically by instrumenting the code to intercept the synchronization primitives. Therefore, the debugging process is simple and applicable without any change in program code by the user. But there are many different situations where a detector is not able to detect the correct semantics of synchronizations between threads or may miss synchronizations. This causes detectors to produce false alarms or miss true data races.

Currently, dynamic race detectors produce many false warnings. The user has to examine all warnings reported by a tool in order to find the true data races. Examining all warnings is a very time-consuming and difficult task. Most of the time, there are only few true races in a program. Thus, the benefit of using an automatic race detector that reduces the number of false warnings is outweighed by manually analyzing the results. For instance, the *Eraser-style* detectors based on the lockset algorithm overwhelm users with many false warnings because of shortcomings in the lockset algorithm. A simple example is provided in Figure 1.5. A detector based on the lockset algorithm considers only *lock/unlock* operations. As a result, it produces a false positive on `DATA`, because the algorithm is not able to identify the existing happens-before relation induced by *signal/wait* primitives.

Furthermore, race detectors miss races (cause false negatives) in various situations. As an example, pure happens-before detectors easily overlook the race on variable `DATA` in Figure 1.6, since the happens-before relation constructed

```
DATA++
signal(CV)
   .
   .
   .
```

(a) Thread 1

```
wait(CV)
   .
   .
   .
DATA--
```

(b) Thread 2

Figure 1.5.: False positive generated by the lockset algorithm.

by *unlock/lock* between Thread 1 and Thread 2 is considered by mistake for the variable DATA too. However DATA is not protected by any lock. Another point is that pure happens-before detectors need a certain amount of access ordering history for each shared variable in order to identify conflicts. This point becomes a serious problem for long-running applications.

```
DATA++

lock(l)
  V = V + 1
unlock(l)
```

(a) Thread 1

```
lock(l)
  V = V + 1
unlock(l)

DATA--
```

(b) Thread 2

Figure 1.6.: Simple example causes a false negative in pure happens-before detectors.

Another difficulty is that race detectors are not able to handle synchronizations involving *condition variables* (*inter-thread event notifications*). Such synchronizations are hard to detect. Also in some cases, it is extremely difficult to construct the implicit ordering imposed by the synchronization primitives used with condition variables. However, inter-thread event notifications are widely used in programs and cause detectors to produce lots of false warnings and even miss races.

Figure 1.7 depicts an example using inter-thread event notifications. Thread 1 operates on DATA and then signals Thread 2 that it can take over the data for further processing. The threads are properly synchronized, but there is an ordering in which the happens-before relation caused by `signal()` and `wait()` is not visible to the race detector. If Thread 1 finishes first, Thread 2 would not call `wait()`. Consequently, the signal sent by Thread 1 is lost. Any instrumentation of `signal()` and `wait()` thus does not detect the proper ordering of the two threads. Thread 2 carries on and as soon as it accesses DATA, a data race is reported, even though there is none. The proper ordering is enforced by the condition variable COND, but noticed by neither lockset nor happens-before detectors.

```
DATA++

lock(l)
  COND = 1
  signal(CV)
unlock(l)
```

(a) Thread 1

```
lock(l)
  while(COND != 1)
    wait(CV)
unlock(l)

DATA--
```

(b) Thread 2

Figure 1.7.: Inter-thread event notifications using synchronization primitives signal() and wait().

Ad-hoc synchronization is a major issue for race detectors. The lack of knowledge of these kind of synchronization operations leads to numerous reports of false warnings. If a detector is aware of all ad-hoc (programmer-defined) synchronizations that occur during program execution, a significant limitation of the race detector is removed. Ad-hoc synchronizations are implemented in program source code itself by the programmer rather than in libraries. All these synchronizations may cause false warnings by a race detector that does not recognize them.

Ad-hoc synchronization may be implemented in various forms, such as *flags* or *spin-locks*. There are different algorithms to accomplish each kind of synchronization operation. Let us consider the simple *flag synchronization* shown in Figure 1.8.

```
DATA++
  .
  .
  .
FLAG = TRUE
```

(a) Thread 1

```
while(FLAG != TRUE){
   /* do_nothing */
}

DATA--
```

(b) Thread 2

Figure 1.8.: Simple ad-hoc synchronization using a FLAG.

Identifying this kind of synchronization operation is not an easy task. When Thread 2 starts executing, it spins on variable FLAG, which will only be modified by Thread 1. Thread 2 cannot proceed, until the shared variable FLAG is set to TRUE by Thread 1. By definition, the executions of the write operation of Thread 1 and the read operation of Thread 2 on variable FLAG represents a dynamic data race. However, the purpose of this data race is only to ensure execution order. Thus, this benign race does not constitute a concurrency bug, if it is intentionally programmed to implement a synchronization construct. Furthermore, a false warning is reported on the shared variable `DATA`, in spite of the fact that both threads are correctly synchronized.

Finally, synchronization primitives from different unknown libraries that are not supported by the detector cause problems. Since knowledge of all synchronization operations in a program is crucial for race detection, missing some primitives means inaccuracy in reports and additional false positives. It is unreasonable to assume that the detector directly supports synchronization primitives from many different libraries. Thus, removing this limitation necessitates a general method to make the detector aware of all synchronization operations happening in a program.

1.3. Structure of the Thesis

This thesis is organized as follows: Chapter 2 explains the objectives. We specify our contribution and set up the hypotheses. In Chapter 3, we define some important terms and present theoretical background. We give a short overview about the lockset based detection algorithm , happens-before detection and some hybrid methods. Chapter 4 discusses past and related work. Chapter 5 contains the new race detection approach. We present our algorithm and the new features used in our dynamic race detector. Different methods to overcome the problems dealing with synchronization by *inter-thread event notifications* and *ad-hoc synchronizations* are described in Chapter 6 and 7.

In Chapter 8, the implementation of our approach is discussed. Furthermore, some important details are depicted. In Chapter 9 our approach is examined with different benchmarks and applications, and the preliminary results are evaluated and compared with other race detectors. In the last Chapter, we give a summary with a short discussion of our results and the focus of our ongoing and future work.

Chapter 2.

Objectives and Contributions

2.1. Objectives

The primary goal of this work is to present a practical and efficient race detection approach to reduce the number of false positives and false negatives (missed races). The approach lessens the number of false alarms so that it pays to analyze the reports. If the number of false warnings in the examined program is small, it is feasible for developers to analyze all of them. However, the approach should not mask races. It should be automatic, easy to use and should not require any manual source code annotations or formal specifications of the program. The races in the program have to be reported with enough contextual information. Providing additional information and guidelines on the suspicious contexts along with the reported warnings makes analyzing the results easier.

2.2. Contribution

The major contribution of my work consists of three parts. The first part is the new race detection algorithm which is adaptable to *long-running* and *short-running* applications. We observed limitations in previous race detection methods and developed a new algorithm to overcome these limitations. Our algorithm has been implemented as part of our new tool[1], which is called Helgrind$^+$ [23, 21]. The user can select the sensitivity depending on her/his preference, or choose different levels of happens-before analysis. We examined our method using substantial benchmark suites. We analyzed the results and compared them with the results of other race detectors.

[1]Helgrind$^+$ is an open source tool and can be downloaded at the following address: http://svn.ipd.uni-karlsruhe.de/trac/helgrindplus

In the second part, we developed a new method for correctly handling synchronization with inter-thread event notifications automatically, without relying on source code annotation. The method accurately establishes happens-before relations implied by condition variables, and thus eliminates almost all cases of false alarms and missed races caused by inter-thread event notifications. We verify our method by implementing it and evaluating it with several benchmarks.

The third part of this work presents a method to detect ad-hoc synchronizations. Identifying the synchronization operations implemented in the program source code itself, and not as synchronization primitives in the libraries, is not trivial. We propose a dynamic software technique that identifies these kinds of synchronization patterns automatically. Such synchronization patterns may happen via flag synchronization or more complex constructs such as barrier synchronizations or spin locks. The proposed method is general and extensible to detect synchronization primitives from unknown libraries that are not supported by the detector. We implement the method in Helgrind$^+$ and confirm that our dynamic technique is able to correctly find ad-hoc synchronizations, and removes false positives without introducing additional false negatives.

2.3. Hypotheses

The ideal dynamic race detector detects all explicit and implicit synchronization operations in order to notify the programmer of incorrect or missing synchronizations in a program. It is aware of all synchronization calls in programs and intercepts them to provide precise reports on unsynchronized accesses which may be a source of bugs. The following requirements are necessary to have a precise and efficient race detection which is able to detect all different kinds of synchronization operations in parallel programs:

1. The whole synchronization process has to take place within the synchronization operation: At the end of calling the synchronization operation, it has to be clear to the detector whether the synchronization operation is successfully finished or not.
2. The detector has to detect all synchronization primitive calls in a program and must be aware of their semantics.
3. Any ad-hoc or implicit synchronization operations along with their semantics must be identified by the detector.

The first requirement relates to *inter-thread event notifications*: They use *condition variables* as a part of their synchronization. Condition variables are implemented in programmer code itself and not in the library. Consequently, inter-thread event notifications build complex synchronization constructs that use a part with synchronization primitives e.g. *signal/wait*-operations for their event notification and another part for some kind of ad-hoc synchronization. By intercepting synchronization primitives from libraries only, the semantics of other part of synchronization in programmer code is missing and hidden to the detector. So, after intercepting the synchronization primitive calls (*signal/wait*), the detector does not know if the synchronization between threads has successfully terminated or not. The need of additional information about the part in programmer code is essential for correct establishment of a happens-before relation.

Our first hypothesis states that a detector has to be aware of the meaning of different synchronization primitives. It has to be able to deal with all of them by a proper algorithm to provide good results.

Hypothesis 1 *By combining the lockset and happens-before analyses, it is possible to correctly handle synchronization primitives, and reduce the number of false positives and false negatives, compared to existing race detectors.*

Thus, we need to explicitly define the semantics of different synchronization primitives for the race detector. Furthermore, we have to provide an efficient race detection algorithm to correctly handle each specific synchronization primitive in order to have a precise race detection.

The second point deals with synchronization primitives that could be specified within a particular library. If the functionality of all synchronization primitives provided in the library is known to the detector, the intention and the exact semantics of the synchronization primitive call are available.

What if synchronization primitives are unknown to the detector or come from unsupported libraries? Then the detector is not able to intercept them. That is, either the program is allowed to call only known synchronization primitives provided in the supported libraries, or the detector must be able to identify unknown primitives from unsupported libraries. Any unknown synchronization primitives used from unsupported libraries have to be identified by a practical race detector. Otherwise, as soon as synchronization primitives are not used and instead of them ad-hoc or unknown synchronizations are used, the detector misses the synchronization and produces false warnings. We set up our second hypothesis based on the above points:

Hypothesis 2 *It is possible to build a universal race detector, that is, a race detector which is not limited to a specific set of libraries, and is able to detect synchronization operations.*

We require a *universal race detector* to identify all synchronization operations in the program which is crucial for a practical race detector to reduce the number of false positives and false negatives.

Chapter 3.

Basic Concepts of Race Detection

In this chapter, we first define some of the terms used later on. Then, we describe basic algorithms used in dynamic race detection followed by a short discussion of their limitations.

3.1. Definitions

Data Races

Data races are synchronization defects in parallel programs. A data race occurs, when two or more threads access a shared memory location which is not protected by a proper synchronization construct (e.g., a lock/unlock), and at least one of them writes the memory location. They are not necessarily considered defects, since they could be intentional. Data races tend to manifest themselves randomly and are troublesome to reproduce and remove.

Data races are caused by *atomicity violations* or *order violations* [27, 24]. We explain them in detail in the sections below.

Atomicity Violation

An *atomicity violation* happens if a *critical section*[1] is interrupted and accessed by more than one thread simultaneously. Atomicity violation could lead to an inconsistent and faulty behavior of a program, which is hard to detect. An example of an atomicity violation is provided in Figure 3.1. Two threads enter

[1]A critical section is an atomic section that has to be executed only by one thread at a time.

an unprotected critical section and increment a shared variable which leads to inconsistent results. Variable COUNTER is a shared variable and TEMP is a thread local variable.

```
TEMP = COUNTER
TEMP = TEMP + 1
COUNTER = TEMP
```

(a) Thread 1

```
TEMP = COUNTER
TEMP = TEMP + 1
COUNTER = TEMP
```

(b) Thread 2

Figure 3.1.: Atomicity violation causes a data race.

Avoiding the atomicity violation is possible by allowing only one thread at a time to enter the critical section. We ensure this by using locks to protect the critical region. Figure 3.2 demonstrates how to get rid of the atomicity violation.

```
lock(l)
  TEMP = COUNTER
  TEMP = TEMP + 1
  COUNTER = TEMP
unlock(l)
```

(a) Thread 1

```
lock(l)
  TEMP = COUNTER
  TEMP = TEMP + 1
  COUNTER = TEMP
unlock(l)
```

(b) Thread 2

Figure 3.2.: Avoiding an atomicity violation by protecting the critical section with locks.

Most of the time, atomicity violations lead to data races as in the previous example. However, atomicity violations may occur without resulting in a data race. For instance, if we alter the previous example as in Figure 3.3, no race happens on COUNTER, but atomicity is violated. All the accesses to COUNTER are protected by a lock with no data race on COUNTER. However, the results are inconsistent because of the atomicity violation. As before, TEMP is a thread local variable.

Finding atomicity violations in a program is not easy for a detector, because the semantics of the program and the intention of programmers might not be recognizable by the detector. For instance, we cannot recognize the critical regions in a program, if we do not have some clue or code annotations from programmers.

```
lock(l)
  TEMP = COUNTER
unlock(l)

TEMP = TEMP + 1

lock(l)
  COUNTER = TEMP
unlock(l)
```

(a) Thread 1

```
lock(l)
  TEMP = COUNTER
unlock(l)

TEMP = TEMP + 1

lock(l)
  COUNTER = TEMP
unlock(l)
```

(b) Thread 2

Figure 3.3.: Atomicity violation within the critical section. But it does not effect a data race.

Order Violation

Order violations happen if program blocks are not executed in the order the programmer expected. Applying a suitable synchronization construct enforces the correct order of execution between the program blocks. Since the developer's intention is not recognizable from the program itself, detecting order violation automatically is difficult. The only possibility is to detect the resulting data race caused by order violations.

The following program depicts an example of order violation. The main thread creates two different threads; the first thread for initializing the data and the second thread for processing the data. The correct interleaving is that the first thread initializes `DATA` before any other thread uses it. But in the actual execution there is no guarantee of the correct interleaving and the second thread may process `DATA` before it is initialized, resulting in a data race on `DATA`.

3.2. Data Race Classifications

Race detectors indicate important hints that are quite helpful to locate a large set of concurrency bugs. We have to distinguish between different categories of data races, and show only the problematic and harmful data races to developers to avoid overwhelming them with too many warnings. We categorize data races as shown in Figure 3.5 and explain them below.

Apparent data races happen, if synchronization is present and the detector doesn't recognize it [40, 44]. The result is false warning. Apparent races may also occur, if the program is switched to a new library and uses synchronization

```
int DATA;

main()
{
  create(&thread_1, &init_data);
  create(&thread_2, &process_data);
}
```

```
void init_data()
{
  DATA = 1;
}
```

(a) Thread 1

```
void process_data()
{
  assert( DATA == 1 );
  do_something(DATA);
}
```

(b) Thread 2

Figure 3.4.: Order violation results in a data race on `DATA`.

primitives that are unknown to the detector, or when some ad-hoc synchronizations are defined by a programmer.

When a *real race* happens, accesses to a given storage location are not synchronized and at least one of them is a write operation. We differentiate between *intentional* and *non-intentional* races. For instance, in some situations a data race is intentionally programmed to implement synchronization constructs (e.g. barriers) and introduce nondeterminism into a program. These intentional races are known as *synchronization races*.

Non-intentional races are not known to programmers and they are not aware of them. The *true races* are *non-intentional* races. However, there are situations where true data races do not effect the program behavior and don't cause any fault. We classify these races as *benign races*, since they do not affect the behavior of the program. As an example, if we want to display the status of a parallel calculation to a user, it is not an issue whether the displayed values are completely consistent. Contrary to *benign races*, *harmful races* could cause an inconsistent program behavior and counted as faulty code.

Ideally, race detectors should only report data races that fall into the *harmful* category. Since detectors try to estimate the semantics of a program at best, apparent data races, intentional and benign data races lead to various false alarms produced by detectors.

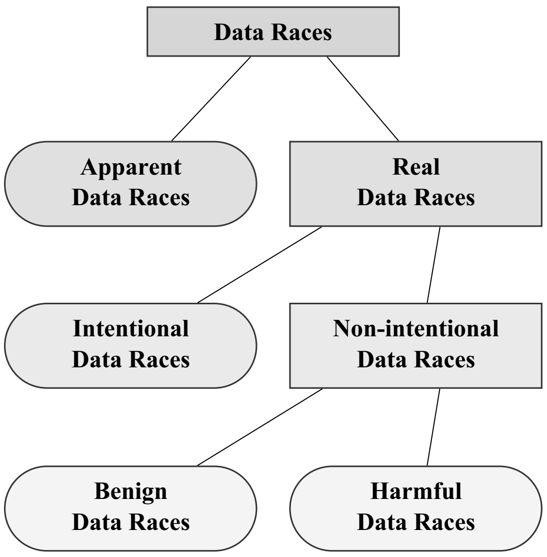


Figure 3.5.: Classification of data races.

3.3. Synchronization

A program protects accesses to data by *synchronization*, avoiding inconsistent program states. The synchronization is guaranteed by providing *mutual exclusion* or temporal ordering of accesses.

Locks implement the principle of mutual exclusion (*mutex*), and assure the atomicity of critical regions. A lock can be attained by one thread at a time. If another thread likewise wants to attain this lock, it will be blocked until the first thread releases the lock.

In this way, locks can protect shared data well, but one should pay attention that the *locking discipline* is not violated. This means that the same data has to be protected by the same lock(s). In Section 3.4.1, we describe the lockset algorithm that is able to check a program if a reliable locking discipline exists.

Additionally, threads could be synchronized by a partial temporal ordering by using synchronization operations such as *inter-thread event notifications* (condition variables), *barriers* or *fork/join*. Thus, a thread could be divided into subsequent segments, called *thread segments* (see 3.4.2.1). All accesses within a thread segment happen before, after or parallel to all other accesses of another thread segment.

A *barrier* is aimed for a group of threads that have to wait until all threads in the group arrive at the barrier boundary. Accesses happenning before the barrier happen before any access after the barrier.

When creating a new thread by a *fork* operation, there is already an implicit temporal ordering between parent and child thread that allows the parent to deliver data to the child without any problem. By *joining*, a thread waits for the termination of another thread.

By using inter-thread event notifications via *condition variables* for the synchronization, threads wait for an arbitrary condition to be fulfilled. If the condition does not come true, the thread will be blocked. As soon as a thread affects the condition, it wakes up the waiting thread. This could be interpreted as sending a signal by a thread to the waiting thread. All the accesses happen by signaling thread before sending the signal, are temporally before all the accesses of waiting thread happen after receiving the signal.

In fact, condition variables cause some influential problems for race detectors – we will discuss them detailed in Section 6.

3.4. Dynamic Data Race Detection

3.4.1. The Lockset Algorithm

The lockset algorithm is based on the observation that each shared memory location accessed by two different threads should be protected by a lock, if at least one access is a write. The detector examines all locations where a shared variable is accessed, and checks whether the shared variable is protected by a lock. If the variable is not protected, a warning is issued. The algorithm is simple and easy to implement. *Eraser* [42] was the first implementation of the lockset algorithm, which worked with programs using the POSIX-Threads library.

In this implementation, *mutex* is the basic synchronization primitive with methods to acquire and release it. A mutex is an object that ensures mutual exclusion on a shared variable. If the mutex is free, a thread acquires it (locks the mutex) and begins to use the shared variable. If however the mutex was already acquired (locked) by another thread, the thread blocks until the thread holding the mutex releases it.

The pseudo code of the basic lockset algorithm or so-called Eraser algorithm [42] is shown in Figure 1. During program execution, the algorithm maintains for each shared variable d a set of locks C_d that contains the intersection of the sets of locks that were held during all accesses to variable *d*. The details of the algorithm appear in [42].

```
Let $L_t$ be the set of locks held by thread $t$.
foreach variable $d$ do
    initialize $C_d$ to the set of all locks
end
On each access to $d$ by thread $t$
    set $C_d \leftarrow C_d \cap L_t$
    if $C_d = \emptyset$ then
        issue warning
    end
end
```

Algorithm 1: Basic lockset algorithm.

The main drawback of the Eraser algorithm is that it produces too many false alarms, because it can only process lock operations, and fails when other synchronization primitives or ad-hoc synchronizations are used. For example, numeric algorithms often consist of several steps separated by barriers. If memory accesses by two separate steps overlap, Eraser would falsely report races, even though they are prevented by the barriers. An algorithm based on the happens-before analysis would not report any false positives in this situation.

A single write operation followed by read-only accesses is a frequent case which lockset detectors must handle. Consider a shared variable that is written once by a main thread and subsequently read by worker threads. It appears that no lock is needed. However, a pure lockset detector would report a race in this case. To handle this situation, Eraser uses the state machine in Figure 3.6. The idea is to defer error reports until a second thread performs a write operation, and reaches the *Shared-Modified* state in the diagram. After allocation, the memory location is in the state *New*. During the first write, it enters state *Exclusive* and leaves this state only if another thread reads or writes the memory location. An error is reported if the state *Shared-Modified* is reached and the lockset is empty.

The example provided in Figure 3.7 includes a data race on the shared variable GLOB. Table 3.1 demonstrates stepwise a possible execution order. It shows how the data race is caught by the lockset algorithm based on the state machine. The race is reported in the state *Shared-Modified* because of the empty lockset.

However, the state machine in Figure 3.6 may mask a race and produce false negatives. The program listed in Figure 3.8 contains a simple undetected data race between main and worker threads. The main thread may write the variable

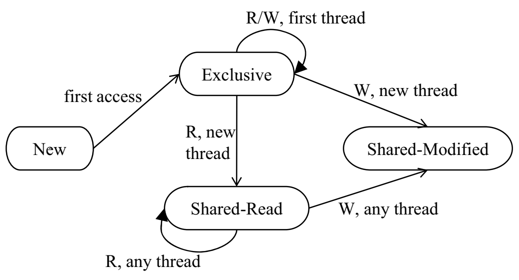


Figure 3.6.: Possible states for a memory location in the basic lockset algorithm.

```
int GLOB = 0;

int main()
{
  create(threadid, worker);

  GLOB = 1;
  printf(GLOB);

  join(threadid, NULL);
}
```

(a) main thread

```
void worker()
{
  lock(l);
    GLOB++;
  unlock(l);

  return NULL;
}
```

(b) worker thread

Figure 3.7.: Detecting a data race on the shared variable `GLOB` by lockset algorithm.

`GLOB` before the worker thread can read it. In this case, the state machine ends up in state *Shared-Read* without issuing a warning. With the opposite order of execution, Eraser would report a race. The basic problem is that there is no synchronization between main and worker threads. As there is no happens-before relation between the main and worker threads regarding the read/write operations, a pure happens-before detector would detect the masked race. This kind of false negatives also exists in other race detectors based on the Eraser state diagram [19, 49].

We developed an new algorithm and extended the state machine such that it handles the above and similar cases correctly. More details are given in Chapter 3.4.

Line No.	GLOB	Old State	New State	L_t	C_d
main(): 1	Initialization	-	New	{}	{l}
main(): 7	Write	New	Exclusive	{}	{}
main(): 8	Read	Exclusive	Exclusive	{}	{}
worker(): 15	Read	Exclusive	Shared-Read	{l}	{}
worker(): 15	Write	Shared-Read	Shared-Modified	{l}	{}

Table 3.1.: Catching a data race on the shared variable GLOB by lockset algorithm (depicted in Figure 3.7).

```
int GLOB = 0;

int main()
{
  create(threadid, worker);
  GLOB = 1;
  join(threadid, NULL);
}
```

(a) main thread

```
void worker()
{
  printf(GLOB);
  return NULL;
}
```

(b) worker thread

Figure 3.8.: A simple example causes false negative in Eraser-based race detectors.

3.4.2. Happens-Before Relation

It is very useful to know the actual time order of events in a parallel program during race detection. Many false positives of the lockset algorithm are avoidable, if we consider the time order between the shared accesses. The happens-before analysis is based on the temporal order of events. The temporal order in a program can be derived from synchronization operations, and results in happens-before relation. By the means of this relation, we describe the partial time order between accesses within a parallel program. Generally, happen-before relation can be represented by two different techniques: *thread segments* or *vector clocks*. We use the acronym hb-relation instead of the happens-before relation for the rest of thesis.

3.4.2.1. Thread Segments

The instruction sequence of a thread can be sliced into a series of pieces, called *thread segments*. Synchronization with other threads (or thread segments) happens at the start or at the end of each thread segment. Of course, all thread segments belong to a specific thread. Within a thread segment, all operations are totally ordered. The thread segments of each thread are also totally ordered. Synchronization defines a partial order of thread segments. If two thread segments are *not* ordered, they may execute in parallel.

Figure 3.9 shows the thread segment diagram of a potential execution of the program depicted in Figure 1.7. Thread 1 sends a signal to Thread 2. Thus the first part of Thread 1 TS_1 happens before the second part of Thread 2, TS_2'. Both Thread 1 and 2 are accessing variable `DATA`. Because of the ordering, there is a hb-relation between TS_1 and TS_2'. They are correctly synchronized and there is no race here.

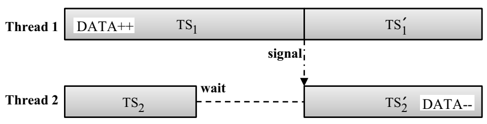


Figure 3.9.: A thread consists of thread segments separated by synchronization operations.

For further discussion, it is useful to define a concise notation for the ordering of thread segments. Lamport's hb-relation $\xrightarrow{hb}$expresses exactly this [25]. When a thread segment TS_1 is executed before another thread segment TS_2, we say $TS_1 \xrightarrow{hb} TS_2$.

We define the relation $\xrightarrow{hb}$ to be reflexive and transitive. The relation is defined to be reflexive regarding thread segments, i.e. $TS_1 \xrightarrow{hb} TS_1$ is possible. This is because execution within a thread segment is strictly ordered and throughout our algorithm, we always compare the present point of execution with a past point of execution which could be in the same segment. Transitivity allows us to traverse through the thread segment graph and check if two segments are parallel: Two thread segments TS_1 and TS_2 are parallel iff there is no path of $\xrightarrow{hb}$-relations between them. This situation is denoted as $TS_1 \parallel TS_2$.

Based on this relation, a potential race has occurred, if we observe that two distinct events are parallel. Compared to lockset-based detection, happens-before analysis has a lower rate of false positives, but causes significant overhead, and is difficult to implement. Moreover, it is sensitive to scheduling.

3.4.2.2. Vector Clocks

Vector clocks are another method to represent the happens-before relation. It is easier and more efficient to implement the happens-before relation by vector clocks. They are based on Lamport clock [25], and assign for each event a global unique time stamp. We used also vector clocks for the implementation of the hb-relation in Helgrind$^+$. However for the sake of simplicity, we indicate the hb-relation by thread segments, when presenting our concept and algorithms.

For each thread a *logical local time* is defined. The logical time is a counter which is incremented by each important event (i.e. synchronization events). The vector clock defines a logical global time and consists of local time of all threads together. That is, a vector clock V is a complete defined function $V : Threads \to \mathbb{N}$.

The hb-relation is defined by vector clocks as the following:

$$V \overset{hb}{\to} W :\Leftrightarrow \forall u \in Threads : V(u) \leq W(u)$$

Each Thread t holds its current time vector V_t. $V_t(t)$ gives the logical local time of the thread, and $V_t(u)$ indicates the most recent local time of thread u which is known to thread t. At the beginning, each thread has no information about the local time of other threads, while its own current local time is initialized to 1. For this reason, the function $newVC$ is defined to produce the initial value of time vectors for thread t:

$$newVC(t) := u \mapsto \begin{cases} 1 & u = t \\ 0 & otherwise \end{cases}$$

When a synchronization happens, the vector clock of the threads is updated. Therefore, we define the two basic operations $join(V, W)$ and $tick(V, t)$ on vector clocks:

$$tick(V, t) := u \mapsto \begin{cases} V(t) + 1 & u = t \\ V(u) & otherwise \end{cases}$$

$$join(V, W) := u \mapsto \max(V(u), W(u))$$

For instance, if *fork/join* happens as shown in Figure 3.10, when a thread creates another thread, the created thread inherits the time vector of parent thread.

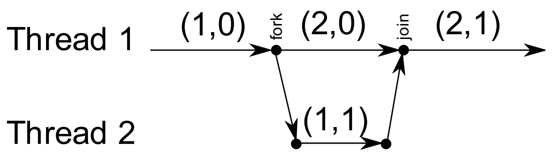


Figure 3.10.: Happens-before relations caused by fork/join operations.

When joining, a thread waits for terminating of another thread. The time vector of terminating thread is taken by waiting thread. The following operations are done when executing fork/join operations:

```
$before\ Thread\ t$ create(u) $executes$:
    $V_u \leftarrow newVC(u)$
    $V_u \leftarrow join(V_t, V_u)$
    $V_t \leftarrow tick(V_t, t)$

$after\ Thread\ t$ join(u) $executes$:
    $V_t \leftarrow join(V_u, V_t)$
```

Similarly, when executing other synchronization operations such as barriers or condition variables, vector clocks are calculated by the help of the above defined functions.

Happens-before analysis uses vector clocks to check if there is a hb-relation between accesses. The following example (Figure 3.11) contains a data race on the shared variable GLOB. The example uses signal/wait primitives for the synchronization which cause the hb-relation depicted in Figure 3.12. The values of vector clocks are also shown. The waiting thread takes the time vector of signaling threads and update its own time vector. So, this relation is valid: the time vector before signaling thread $\xrightarrow{hb}$the time vector after the waiting thread, i.e. $(1, 0) \leq (1, 1)$. At the point where the race happens (line 16), time vectors do not indicate any hb-relation and threads are in parallel (as shown in Table 3.2).

Line No.	**GLOB**	$\xrightarrow{hb}$	$V_{Thread1}$	$V_{Thread2}$
main(): 1	Initialization	-	(1,0)	(0,1)
main(): 8	Write	-	(2,0)	(0,1)
worker(): 29	Write	yes	(2,0)	(1,1)
main(): 16	Write	no	(2,0)	(1,1)
main(): 17	Read	no	(2,0)	(1,1)

Table 3.2.: Catching a data race on the shared variable GLOB by happens-before analysis (depicted in Figure 3.11).

```
int GLOB = 0;
int COND = 0;

int main()
{
  create(threadid, worker);

  GLOB = 1;
  sleep(2000);

  lock(l);
    COND++;
    signal(cv);
  unlock(l);

  GLOB = 3;
  printf(GLOB);

  join(threadid, NULL);
}
```

(a) Thread 1

```
void worker()
{
  lock(l);
  while(COND !=1){
    wait(cv);
  }
  unlock(l);

  GLOB = 2;

  return NULL;
}
```

(b) Thread 2

Figure 3.11.: Detecting a data race on the shared variable GLOB by happens-before analysis.

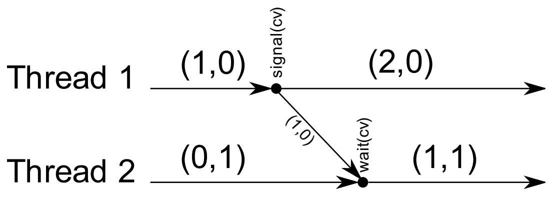


Figure 3.12.: Happens-before relations caused by signal/wait operations.

Chapter 4.

Related Work

In this chapter, we provide an overview about the previous work in the area of race detection techniques in parallel programs. There is a substantial amount of prior work regarding detection of potential data races. Proposed solutions can be roughly classified as static (ahead-of-time) and dynamic (on-the-fly) analyses. We present and discuss the advantages and disadvantages of these methods. Finally, we talk about some further techniques for race detection e.g. software transactional memory.

4.1. Static Analysis

Static analysis considers the entire program and warns about possible races caused by all possible execution orders [14]. The main drawback of this approach is that it produces many false positives, as static analysis conservatively considers all potential thread interleavings, even those that are not feasible. Another issue is that static analysis does not scale well to large programs due to state space and path explosion problems [10]. Furthermore, static analysis has problems with dynamically allocated data, since it has no information about it. Detecting all feasible data races by static analysis is known to be an NP-hard problem [34]. For this reasons, most current static race detectors (e.g. [34]) focus on identifying a subset of data races.

Some static techniques are based on strong type-checking and assume that well-typed programs are guaranteed to be free of data races [16, 5]. They introduce a new static type system for multi-threaded programs to prevent data races. In fact, the new type system allows programmers to specify the locking discipline in their programs in the form of type declarations. They use ownership types to prevent data races. Ownership types provide a statically enforceable way of specifying object encapsulation. This method is limited to a specific language and requires type annotations, either inferred by the type systems or manually annotated by programmers.

Compilers widely use control flow and data flow analyses to optimize programs. Both of these techniques are also applicable to race detection during static analysis. Data flow analysis identifies program invariants at each program point by propagating information along control flow paths. The control flow graph for parallel programs is the combination of the control flow graphs of the individual tasks. For parallel programming models with shared memory, every instruction of an arbitrary task can be a direct successive control flow block for a given instruction. This could lead to path explosion in the graph; flow analysis would take a long time [8]. Reducing the number of paths in the parallel control flow graph is done by identifying synchronization constructs that prevent parallel execution and remove some infeasible edges between the tasks.

The lockset algorithm is also used in a static tool called *RacerX* [14]. It uses flow sensitive, inter-procedural analysis to detect race conditions. It checks information such as which locks protect which operations, which code contexts are multi-threaded, and which shared accesses are malicious. RacerX examines system-specific locking functions by extracting a control-flow graph from the system, which is used for further analysis to find races. The tool has performance problems and lacks a reasonable pointer analysis. It has only simple function pointer resolution.

4.2. Model Checking

Another static analysis method is *model checking*. Model checking is used to statically verify program properties specified in temporal logic. The timing behavior of a concurrent program is expressed in temporal logic to find concurrency bugs. For instance, Java PathFinder (JPF) [20] is able to find violations of any assertions written in Java, such as race conditions. Generally, model checking does not scale well, and without appropriate abstractions even JPF is only applicable to programs smaller than 10000 lines of code (10 KLOC). This is because of the state-space explosion, as in other static methods.

In addition to the state explosion problem, model checking suffers from high initial overhead due to the need for annotations. Thus, having an algorithm that extracts all needed information for analysis directly form the source code is more desirable.

4.3. Dynamic Analysis

Dynamic analysis scales better and reports fewer false positives compared to static analysis. However, it detects races only in actual executions. Consequently, the program has to be tested with various inputs to cover different execution paths and interleavings. There are two different methods used by dynamic race detectors: on-the fly and post-mortem. On-the-fly methods record and analyze information as efficiently as possible during program execution. Post-mortem methods record events during program execution and analyze them later. All dynamic methods add overhead at runtime, which must be traded off against detection accuracy.

Prior dynamic race detectors are based on two different techniques: lockset or happens-before analysis. Lockset analysis checks whether two threads accessing a shared memory location hold a common lock. If this is not the case, the concurrent access is considered a potential data race [42, 48, 35]. The technique is simple and was introduced for the first time in Eraser [42]. The lockset algorithm can be implemented with low overhead and is relatively insensitive to execution order. The main drawback of a pure lockset-based detector is that it produces many false alarms due to the fact that it ignores synchronization primitives other than locks, such as signal/wait, fork/join, and barriers.

Dynamic detectors [48, 35] use escape analysis to determine if variables are used by only a single thread. They filter out these variables and non data race statements to reduce the runtime overhead. In addition, they detect data races at object level instead of at the level of each memory location. Object-oriented languages such as Java allow users to restrict access to structures at compile-time. They have extended the ownership model of Eraser such that a transfer of ownership is allowed once for every object. They carry out the expensive lockset operations only for the shared objects. However, it is difficult to apply dynamic escape analysis to languages that can access any memory location through pointers, such as C/C++.

Happens-before detectors [47, 9, 6] are based on Lamport's happens-before relation [25]. Happens-before analysis uses program statement order and synchronization events to establish a partial temporal ordering of program statements. A potential race is detected if two threads access a shared memory location and the accesses are temporally unordered. The happens-before technique can be applied to all synchronization primitives, including signal/wait, fork/join, barriers, and others. It does not report false positives in the absence of real data races. However, this approach may miss races, i.e. produce false negatives, as it is sensitive to the order of execution and depends on whether or not the scheduler generates a harmful schedule. In fact, happens-before detection produces more false negatives than lockset-based detection [36]. Happens-before analysis is also difficult to implement efficiently and does not scale well.

Combining the happens-before analysis with lockset analysis results in a hybrid solution with a trade-off between accuracy and runtime overhead. Recent race detectors [36, 19, 41, 49, 39, 45] have combined happens-before and lockset-based techniques to get the advantages of both approaches. The combination was originally suggested by Dinning and Schonberg [12]. However, combined approaches, which are referenced above, still produce many false positives and miss races. Additionally, they are limited to a particular library and support only a subset of synchronization primitives.

Hybrid detectors extend the Eraser algorithm by using happens-before analysis. The constructs fork/join and an elementary interception of signal/wait on condition variables are regarded to establish hb-relations. Unlike our approach, they use signal/wait directly to establish hb-relations, which is not valid in cases with lost signals and spurious wake ups for condition variables. In cases of lost signals, they don't construct hb-relations and for spurious wake ups, they create false hb-relations. Visual Thread and Helgrind [19, 45] build hb-relations only on fork/join operations. They consider memory locations that are limited to non-overlapping thread segments as exclusive even if they are shared, and not accessed by a single thread.

A few publications [49, 39] use techniques for adaptive granularity and apply variable size detection units. Choosing a small detection unit might result in higher overhead, while choosing a large one might lead to false positives. RaceTrack [49] switches the object to field granularity during race detection. It starts with object level race detection and automatically refines its algorithm to field-level detection, when a potential race is detected. Generally, object-size granularity has limitations such as performance considerations and it suffers from the inability to correctly detect object boundaries.

Most of the detectors above use vector clocks for the happens-before analysis to track the order of events. Using vector clocks to track hb-relations requires history information. Vector clocks are easier to implement compared to thread segments. The approaches in [19, 45] use thread segment, and need less memory. They work for synchronization primitives other than fork/join, too.

Dynamic race detectors differ in how they monitor program execution. Many detectors [42, 19, 45, 47] use binary instrumentation. They record load/store instructions, references to shared-memory locations and synchronization primitives from binary files and instrument them. The race detector Helgrind$^+$ implemented in this work falls into this category. Other race detectors [49, 36, 9] work with the bytecodes of object-oriented programming languages, making them independent of programming language and source code. Some race detectors such as MultiRace [39] modify the source code in order to instrument memory accesses and synchronization instructions. The Intel Thread Checker [41] instruments either source code or binary code of the program.

Our dynamic detection approach is also based on the lockset algorithm and happens-before analysis, but the heuristics employed and the combination of these methods differentiate our detector from other approaches. We propose two new memory state models that are optimized for short-running and long-running programs [23, 21]. Compared to the simple memory state models presented in previous papers [49, 39, 36, 45, 41, 19], our models take full advantage of happens-before analysis and the accurate detection of happens-before relations.

Furthermore, the above mentioned detectors [49, 39, 36, 45, 41, 19] produce a lot of false warnings and even miss races. This is because the ordering induced by inter-thread event notifications and ad-hoc synchronizations are not taken into account. Basically, they suffer from two serious limitations: (a) they are not able to detect ad-hoc synchronizations implemented in user code; (b) the detectors are restricted to synchronization primitives of a specific library,– synchronization primitives from other libraries are ignored. Thus, they are not able to produce accurate reports, and applying these detectors to real applications overwhelms the user with too many false alarms. Our work removes these limitations by introducing a general approach for detecting ad-hoc synchronizations and unknown synchronization primitives. We are able to eliminate false positives including benign synchronization races and possible false negatives caused by missed or incorrect synchronizations.

Tian et al [44] used a dynamic technique to identify synchronization operations. The technique is able to partially suppress false positives caused by apparent races and benign synchronization races. It is based on the actual spinning reads, which may occur at runtime, and sets a threshold value for the number of spinning reads to identify them during execution. The value of the threshold is set heuristically (they set the number of spin reads to three). If the spinning read does not happen, the detector will not construct a hb-relation and miss the synchronization. This could happen when using condition variables or ad-hoc synchronization that causes false alarms. Furthermore, the method could, by mistake, identify ordinary loops in the program as spinning reads and interpret them as synchronization operations. This may lead to misinterpretation and creating false synchronizations by the detector causing false negatives (missed races).

By contrast, our dynamic method for spinning read detection is general and could be used as a complete race detection approach in a race detector. It exploits the semantics and dynamic information of program code to identify ad-hoc synchronizations along with different synchronization operations. The resulting race detector is a universal happens-before race detector. Compared to other happens-before race detectors such as DRD 3.4.1 [47], this method also induces happens-before edges when using ad-hoc synchronization or unknown synchronization primitives, resulting in substantial accuracy.

Recently, FastTrack [15] introduced a technique to implement vector clocks in a lightweight manner. The technique replaces heavyweight vector clocks with an adaptive lightweight representation. It requires constant space and supports constant-time operations for some operations of the target program. The new representation of vector clocks improves time and space performance during happens-before analysis. Also, another technique introduced by Goldilocks [13] tries to make precise the lockset algorithm by defining new lockset update rules. The new lockset update rules allow a variable's locksets to grow during the execution. In fact, the lockset of a variable may be modified even without the variable being accessed. This technique helps to deal with different cases such as shared data protected by different locks at different points in time, and data protected indirectly by locks on container objects. Both of the introduced techniques can be used by our hybrid race detection algorithm, which could improve the performance of the happens-before analysis and the precision of the lockset algorithm.

4.4. Post-mortem

Post-mortem methods log events during program execution and analyze them after program execution. They are unsuitable for long-running applications, specially if they interact with their environment extensively. While post-mortem analyses can affect performance less than dynamic analyses, they suffer from the same limitation as dynamic techniques, in that they can only find errors along executed paths. Some Post-mortem techniques combine static and dynamic techniques. They collect information at compile time and then analyze the re-execution of the program based on the collected information.

For instance, a system which uses post-mortem techniques for debugging of non-deterministic parallel programs is RecPlay [40]. It records only critical events and then replays them. That is, it records only synchronization operations, and checks for data races using happens-before analysis during a replayed execution. Recplay uses a classical logical vector clock for detecting concurrent accesses.

It is more suitable for programming languages with unconstrained life time and access to shared variables.

Another record/replay tool is *Déjà Vu* for Java programs, which was presented by Choi and Srinivasan [7]. It provides deterministic replay of a program's execution. It introduces the concept of logical thread schedule, that is, a sequence of intervals of critical events. Each interval corresponds to the critical and non-critical events executed consecutively in a specific thread. Déjà Vu records only critical events, namely all synchronization events and the shared variable accesses, by capturing logical thread schedule intervals. At replay time, Déjà Vu reads the thread schedule information from a file created at the end of the recording.

Déjà Vu can only deterministically replay the non-deterministic execution behavior due to thread and related concurrent constructs such as synchronization primitives. However, other events such as input/outputs, window events, and system calls have not been taken into account. This is because replaying a faithful execution requires recording as many non-deterministic events as possible and sometimes may be infeasible and incur intolerable overhead during the recording phase.

4.5. Software Transactional Memory

Software Transactional memory (STM) has been proposed to simplify parallel programming by removing the need for explicit locks [20]. The idea is to use transactions for accessing shared data structures and to use roll backs in case of conflicts. A transaction is a sequence of operations that either commits completely or has no effect. In fact, STM provides an alternate concurrency control mechanism by supporting atomic and isolated execution of programmer-specified tasks. The main limitation with STM is that some operations are not reversible as in database systems. A software transaction could contain I/O-operations which are not retreatable. Another problem is the lack of a mature programming language for STM.

4.6. Hardware Transactional Memory

All of the race detection approaches discussed above are software-based. Usually data race detection is done entirely in software. Recently, RaceTM [18] presents a approach to data race detection by hardware transactional memory. There are few hardware-assisted dynamic race detection mechanisms such as HARD [50] that use lockset-based hardware-assisted race detection and are faster. But they require also specialized hardware, which is not cost-effective. RaceTM leverages transactional memory hardware to perform efficient dynamic data race detection. It demonstrates how the ordinary multicore hardware that supports code with transactional critical sections can be used for detecting data races in an entire program, while avoiding the costs of building specialized hardware.

Chapter 5.

Helgrind$^+$ Race Detection

The race detection approach in Helgrind$^+$ is a new hybrid approach that exploits the advantages of combining happens-before analysis and the lockset algorithm to get more accurate results. The presented algorithm overcomes the observed limitations in earlier dynamic tools and proposes two new memory state models which are optimized for *long-running* and *short-running* applications. It takes advantage of more detailed state machines and high accuracy in detecting happens-before relations. We describe our approach in detail in the following sections. In the upcoming chapters, we extend our algorithm for the primitives of inter-thread event notifications(signal/wait). Then, we complete the algorithm by presenting a method for detecting programmer-defined ad-hoc synchronizations.

5.1. The Algorithm

Our detection algorithm combines the happens-before analysis and the lockset algorithm in a novel and efficient way. Basically, both lockset and happens-before analyses are performed. The lockset algorithm is a reliable method to detect correct synchronization using critical sections. However, it ignores the hb-relations caused by lock primitives between locking and unlocking of a critical section. This is similar to previous hybrid race detectors [49, 39, 36]. The happens-before relations derived from other synchronization primitives are taken as usual.

The algorithm takes synchronization primitives into account. Compared to other detectors [49, 39, 36], which consider only a subset of synchronization primitives, we are able to handle all common synchronization primitives. The algorithm can be implemented for synchronization primitives provided in any library. In general, our algorithm considers the points below:

1. The program to be tested uses the following typical synchronization primitives:

   - locks
   - fork/join
   - signal/wait (condition variables)
   - barriers

2. Throughout the program, each shared variable can be protected by different types of the aforementioned primitives. For example, a variable `X` at the beginning of the program could be protected by a lock, whereas later on, a barrier could be employed.

Helgrind$^+$ uses dynamic instrumentation to track program execution. It instruments and monitors every access to memory. To detect races, each variable has an associated state. This state indicates whether the variable is shared or exclusively accessed and which thread segments access it. Every access to a variable is instrumented to track the associated state according to a finite state machine. This state machine is a fundamental part of the race detector. Section 5.2 gives detailed information on the state machine.

The following subsections describe in detail which synchronization operations are instrumented for our algorithm, and how the current state of the program and its threads are maintained. Firstly, we talk about instrumentation of lock operations for the lockset algorithm. Secondly, we describe instrumentation for the happens-before analysis of other primitives, and lastly, we complete our algorithm and combine both methods by presenting the memory state machines.

5.1.1. Lock Operations

Lock operations are handled by the lockset algorithm. We explained the lockset algorithm in Section 3.4.1, which examines if a shared variable is consistently protected by locks. For this reason, we need to know which locks are being held by each thread at any time. The locks held by Thread t are stored in the lockset L_t. When Thread t acquires or releases a lock, we have to update L_t in the following way:

```
After t executes lock(l):                    After t executes unlock(l):
    $L_t \leftarrow L_t \cup \{l\}$                        $L_t \leftarrow L_t \setminus \{l\}$
end                                          end
```

By executing the primitives *lock*(l), l is added to the lockset L_t. Similarly by executing *unlock*(l), l is removed from L_t. In this way, we track the lockset of each thread, and make use of them later on in the algorithm.

5.1.2. Happens-Before Analysis

For the happens-before analysis, the algorithm maintains thread segments and the hb-relations between them. For convenience, we describe the algorithm using thread segment identifiers (TS_t). A notation similar to TS_t could be used for vector clock values. We define the function $NewSegment(TS_1, TS_2, ..., TS_n)$ that performs the following actions: It returns a new thread segment TS_{new} and adds new hb-relations such that $\forall i : TS_i \overset{hb}{\to} TS_{new}$. At any point in time, each thread t is in one of the thread segments. The current thread segment of thread t is called TS_t. When a thread executes a synchronization primitive, the current thread segment ends and a new one is created.

$fork()$ / $join()$ are used for creation and termination of threads. When a thread t creates a new thread u, everything u does happens after t's past operations. Thread u cannot hold any locks at that moment, so L_u is set to empty.

```
Before $t$ executes fork$(u)$:
    $L_u \leftarrow \emptyset$
    $TS_u \leftarrow NewSegment(TS_u, TS_t)$
    $TS_t \leftarrow NewSegment(TS_t)$
end
```

```
After $t$ executes join$(u)$:
    foreach shared variable $d$ do
        $S_d \leftarrow S_d \setminus \{u\}$
        if $S_d$ is singleton then
            reset $d$ to exclusive state
        end
    end
    $TS_t \leftarrow NewSegment(TS_t)$
end
```

When thread t calls $join(u)$, it will wait for thread u to terminate. Everything thread u has done happens before any operation t will do after joining. Additionally, on a $join()$ operation, we scan through all shared variables to see if some of them are no longer shared. Each shared variable d is accessed by a *set of threads* called S_d. If S_d contains a single element after the terminated thread u was excluded from the set, variable d can be reset to the non-shared or "exclusive“ state.

signal() / *wait*() are the primitives for synchronization with *inter-thread event notifications.* A thread t sends a signal while another thread u blocks until a signal is received. Operations of thread t, before sending the signals, happen before operations of thread u, after receiving it. The thread segment of the signaling thread has to be stored so that the waiting thread can create a hb-relation. As different signals can be sent, depending on which condition variable *cv* is used, each condition variable can hold a thread segment called TS_{cv}.

```
Before t executes signal(cv):                After u executes wait(cv):
    $TS_{cv} \leftarrow TS_t$                    $TS_u \leftarrow NewSegment(TS_u, TS_{cv})$
    $TS_t \leftarrow NewSegment(TS_t)$       end
end
```

When using the *barrier*() primitive, each thread is allowed to leave the barrier only after all participating threads have reached it. Thus, each thread segment which is after the barrier happens after all other thread segments before the barrier. A barrier stores an immediate thread segment TS_b. After all participating threads have reached the barrier, TS_b happens after all thread segments. By leaving the barrier, each thread segment synchronizes with TS_b.

```
Before t executes barrier(b):                     After t executes barrier(b):
    $TS_b \leftarrow NewSegment(TS_b, TS_t)$          $TS_t \leftarrow NewSegment(TS_t, TS_b)$
end                                               end
```

5.2. Memory State Machines

The effect of a memory state machine on the outcome of a detector is crucial. With Helgrind$^+$, one can choose between two different memory state machines. Based on our empirical studies, the memory state machines are tailored and carefully tuned for two different categories of applications: long-running and short-running applications. Compared to the memory state machine of Eraser [42] and similar tools, our memory state machines are more complex and accurate. We address the limitations observed in earlier memory states by making the required refinements for lockset and threadset.

We provide both versions of memory state machines (or shorty MSM) in Helgrind$^+$ to have a complete solution for different applications. The user is able to choose the memory state machine depending on the application type.

5.2.1. Memory State Machine for Long-running Applications

The memory state machine tailored to long-running applications [23] is based on the assumption that a past data race access pattern is likely to be repeated in the future. We refer to this state machine as MSM-long. MSM-long has eight different states, and it defers the happens-before analysis until the lockset analysis detects enough insufficiencies. Our empirical results from an analysis of MSM-long showed a significant reduction of false positives[23], making the tools practical for long-running applications.

Figure 5.1 shows the extended memory state machine. A description of each state and the related instructions of the detection algorithm follows in the sections below.

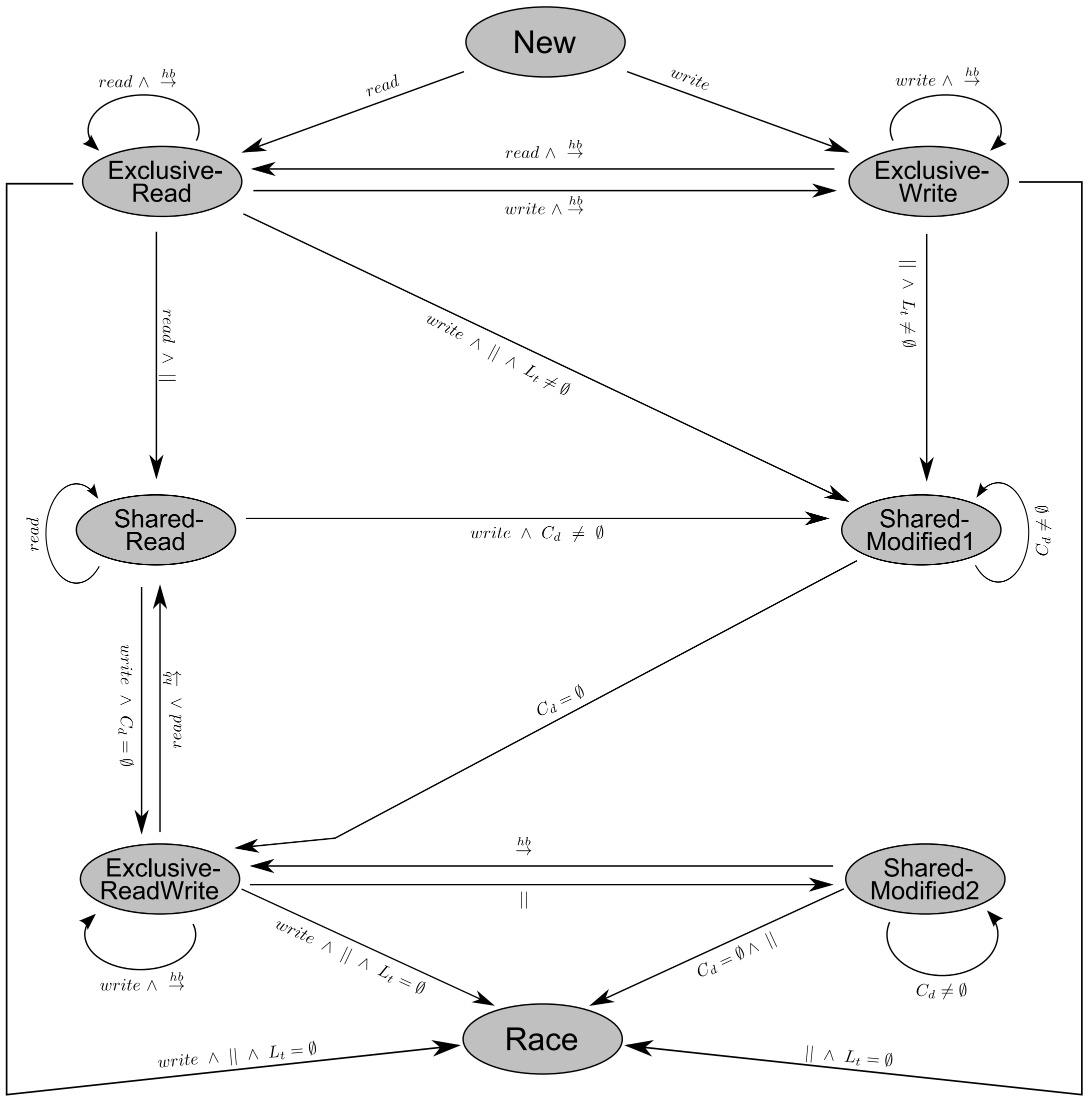

New
read
write
read ∧ $\overset{hb}{\rightarrow}$
write ∧ $\overset{hb}{\rightarrow}$
Exclusive-Read
Exclusive-Write
read ∧ $\overset{hb}{\rightarrow}$
write ∧ $\overset{hb}{\rightarrow}$
read ∧ ||
write ∧ || ∧ $L_t \neq \emptyset$
|| ∧ $L_t \neq \emptyset$
read
Shared-Read
write ∧ $C_d \neq \emptyset$
Shared-Modified1
$C_d \neq \emptyset$
write ∧ $C_d = \emptyset$
read ∧ $\overset{hb}{\rightarrow}$
$C_d = \emptyset$
Exclusive-ReadWrite
$\overset{hb}{\rightarrow}$
||
Shared-Modified2
write ∧ $\overset{hb}{\rightarrow}$
write ∧ || ∧ $L_t = \emptyset$
$C_d = \emptyset$ ∧ ||
$C_d \neq \emptyset$
Race
write ∧ || ∧ $L_t = \emptyset$
|| ∧ $L_t = \emptyset$

Figure 5.1.: Memory state machine for long-running applications.

A *shared* state of a variable is defined, when more than one thread access the variable. An *exclusive* state means that only one thread accesses the variable, or there are hb-relations between successive accesses (accesses are synchronized). In this case, the exclusive ownership of a variable is transferred from one thread segment to another one.

We used the following notation in the MSM diagrams and also for describing the algorithm and different states:

d	an arbitrary memory location.
TS_t	thread segment of the current thread t accessing d.
TS_d	thread segment of the prior thread accessing d.
$\|\|$	current access and the prior access to d are in parallel.
$\overset{hb}{\rightarrow}$	a hb-relation exists between the current and prior access to d.
L_t	current set of locks held by thread t.
C_d	current candidate set of locks protecting variable d.
S_d	current set of threads accessing variable d.

5.2.2. Principles of MSM-long

The general idea is to avoid entering a shared state until the happens-before analysis shows that there is a concurrent access to a memory location. Lockset tracking is performed only in shared states; C_d is updated only in *Shared-Read*, *Shared-Modified1*, and *Shared-Modified2*. No happens-before test is done in the states *Shared-Read* and *Shared-Modified1*. Only state *Shared-Modified2* requires both lockset updates and happens-before analysis (see Figure 5.1). Happens-before analysis is deferred until the lockset of a location is empty, leading to performance improvement. Updating both locksets and the thread segment graph for each access during program execution can be quite expensive in both time and memory consumption. The idea of deferring the computation of happens-before until necessary was introduced for the first time in [39]. This idea is implemented here by introducing the state *Exclusive-ReadWrite*.

Separate *Exclusive-Read* and *Exclusive-Write* states are beneficial for several reasons. As described previously, the Eraser algorithm is vulnerable to scheduling (see the example in Figure 3.8). By introducing these two exclusive states, the state machine can distinguish a read after a write and a write after a read, so the race is detected regardless of schedule, causing an immediate transition to *Race*.

In addition, the edge from state *New* to *Exclusive-Read* makes the detector work more precisely, and handle more cases properly. It is often possible for locations to change from *New* directly to *Exclusive-Read*, especially if the application is reading uninitialized variables, or due to compiler optimizations, where the compiler loads a word from memory, part of which is uninitialized, and then does not use the uninitialized part. Another case is when a program has its own memory management routines that initializes memory with zeros before allocation. In this situation, the memory will be *New* but a read from it is quite legal.

With the edges from *Exclusive-Write*, *Exclusive-Read* and *Exclusive-ReadWrite* to *Race*, we capture races that happen only once at initialization time. An Eraser-style detector is based on the assumption that the program runs for a long time, and if the race happens many times, it will be caught eventually. With the additional edges, the extended memory state machine can catch the race even if it happens at initialization time.

Finally, the idea of introducing a separate state *Race* is useful, because once the race is reported, the tool does not spend time on this memory location any more.

5.2.3. States of MSM-long

For better understanding, we provide a scenario in Figure 5.2 to describe different possible states in the diagram. The variable `GLOB` is a shared variable which is initialized to zero and accessed by threads. We trace the algorithm on this variable during program execution and illustrate state transitions on `GLOB` as shown in Table 5.1.

As before for describing the algorithm, thread segment identifiers (TS_t) are used to indicate a hb-relation between two successive accesses to a memory location (Details about TS_t are discussed in Section 3.4.2). The function $threadof(TS_t)$ returns the thread to which the thread segment TS_t belongs. We use this function when updating threadset S_d.

By executing `GLOB = 0` at line 1 in the given example (Figure 5.2), initially the variable `GLOB` has the state *New*.

```
 1 int GLOB = 0;
 2 int COND = 0;
 3
 4 int main(){
 5
 6   create(threadid, worker);
 7
 8   GLOB = 1;
 9   printf(GLOB);
10
11   lock(l);
12   while(COND !=1){
13     wait(cv);
14   }
15   unlock(l);
16
17   lock(l);
18     GLOB++;
19   unlock(l);
20
21   GLOB++;
22
23   join(threadid, NULL);
24 }
```

(a) main thread

```
25 void worker(){
26
27   sleep (2000);
28
29   printf(GLOB);
30
31   lock(l);
32     COND++,
33     signal(cv);
34   unlock(l);
35   sleep (4000);
36
37   GLOB = 4;
38   printf(GLOB);
39
40   return NULL;
41 }
```

(b) worker thread

Figure 5.2.: A data race occurs on the shared variable GLOB after *two* unsynchronized accesses - overall *nine* accesses.

New: Newly allocated location that is not yet accessed. No lockset is needed. On the first *write/read* access, we enter state *Exclusive-Write/Exclusive-Read.*

```
State New:
  When t executes read(d) ∨ write(d):
    $TS_d \leftarrow TS_t$
    set state to Exclusive-Read / Exclusive-Write
  end
```

After the first write operation on variable GLOB at line 8, we enter *Exclusive-Write*. Only happens-before analysis is performed within this state.

Exclusive-Write: Location d is exclusively written by a particular thread. We track the thread segments and no lockset is needed in this state. As long as write accesses occur from the same thread, we stay in this state (the

Line No.	GLOB	Old State	New State	$\overset{hb}{\rightarrow}$	L_t	C_d
main(): 1	Initialization	-	New	-	-	-
main(): 8	Write	New	Excl-W	-	{}	-
main(): 9	Read	Excl-W	Excl-R	yes	{}	-
worker(): 29	Read	Excl-R	Sh-R	no	{}	{}
main(): 18	Read	Sh-R	Sh-R	-	{l}	{l}
main(): 18	Write	Sh-R	Sh-Mod1	-	{l}	{l}
main(): 21	Read	Sh-Mod1	Excl-RW	-	{}	{}
main(): 21	Write	Excl-RW	Excl-RW	yes	{}	-
worker(): 37	Write	Excl-RW	Race	no	{}	-
worker(): 38	Read	Race	Race	-	-	-

Table 5.1.: Catching a data race on variable GLOB after *two* unsynchronized accesses by MSM-long (depicted in Figure 5.2).

hb-relation holds within the same thread, so there is no concurrent access). We remain in this state as long as each successive write satisfies $TS_d \overset{hb}{\rightarrow} TS_t$, since there are no concurrent accesses to d (TS_d is the thread segment of the prior thread accessed to d). When a write or read occurs with $TS_d \parallel TS_t$, we enter *Shared-Modified1*. In *Shared-Modified1*, we do not track thread segments, so there is no need to store TS_t by leaving *Exclusive-Write*.

```
State Exclusive-Write:
When $t$ executes Read(d) $\vee$ Write(d):
    if $TS_d \overset{hb}{\rightarrow} TS_t$ then
        $TS_d \leftarrow TS_t$
        set state to Exclusive-Read / Exclusive-Write
    end
    else if $TS_d \parallel TS_t \wedge L_t \neq \emptyset$ then
        $C_d \leftarrow L_t$
        $S_d \leftarrow \{t,\ threadof(TS_d)\}$
        set state to Shared-Modified1
    end
    else
        set state to Race
    end
end
```

Variable GLOB is read by the same thread at line 9. By reading the value of GLOB, there is a transition to *Exclusive-Read.* Then, the main thread is blocked by calling the **wait** function at line 13. But the worker thread created by the main thread is running.

Exclusive-Read: Location d is exclusively read by a particular thread. Similar to *Exclusive-Write*, when an access occurs that would result in TS$_d$ || TS_t, we enter *Shared-Read* or *Shared-Modified1.* The information for hb-relation is discarded, as it is no longer needed.

State Exclusive-Read:

```
When $t$ executes read(d):
    if $TS_d \parallel TS_t$ then
        $C_d \leftarrow L_t$
        $S_d \leftarrow \{t,\ threadof(TS_d)\}$
        set state to Shared-Read
    end
    else
        set state to Exclusive-Read
    end
end

When $t$ executes write(d):
    if $TS_d \overset{hb}{\rightarrow} TS_t$ then
        $TS_d \leftarrow TS_t$
        set state to Exclusive-Write
    end
    else if $TS_d \parallel TS_t \wedge L_t \neq \emptyset$ then
        $C_d \leftarrow L_t$
        $S_d \leftarrow \{t,\ threadof(TS_d)\}$
        set state to Shared-Modified1
    end
    else
        set state to Race
    end
end
```

The worker thread accesses GLOB and read its value (line 29). There is no hb-relation between threads. Thus, the current read access on GLOB is in parallel with the last one, and we enter *Shared-Read.*

Shared-Read: Location d is concurrently accessed by multiple threads, but all accesses are reads. We enter this state from *Exclusive-Read* when a read results in multiple concurrent accesses. In this state, we track only the lockset C$_d$, which is initialized to L$_t$. The lockset is updated for every access. If C$_d$ is empty and a write operation occurs, we enter *Exclusive-ReadWrite* to perform happens-before analysis and see whether further accesses are in parallel or not. If C$_d$ is not empty and a write occurs, we enter *Shared-Modified1.* No errors are reported in this state.

```
State Shared-Read:
When t executes read(d):
    $C_d \leftarrow C_d \cap L_t$
    $S_d \leftarrow S_d \cup \{t\}$
    set state to Shared-Read
end

When t executes write(d):
    $C_d \leftarrow C_d \cap L_t$
    $S_d \leftarrow S_d \cup \{t\}$
    if $C_d = \emptyset$ then
        set state to Exclusive-ReadWrite
    end
    else
        set state to Shared-Modified1
    end
end
```

After receiving a signal from worker thread, the main thread continues to work. Variable GLOB is incremented by the main thread at line 18. The increment operation is not atomic and consists of two operations: a read and a write. Reading the value of GLOB does not cause a transition and we remain in *Shared-Read*. After writing the incremented value by the same thread, we enter *Shared-Modified1*, since the main thread holds a lock and C_d is not empty.

Shared-Modified1: Location d is read and written concurrently by multiple threads. This state is entered either from *Exclusive-Write* or *Exclusive-Read*, with C_d initialized to L_t. As in *Shared-Read*, we track only the lockset in this state. If C_d becomes empty, we enter *Exclusive-ReadWrite* to check if there is a hb-relation between the accesses.

```
State Shared-Modified1:
When t executes read(d) ∨ write(d):
    $C_d \leftarrow C_d \cap L_t$ if $C_d = \emptyset$ then
        set state to
        Exclusive-ReadWrite
    end
    else
        set state to Shared-Modified1
    end
end
```

Variable GLOB has the state *Shared-Modified1* before executing the instruction at line 21. After the read access by the main thread as a part of increment, the state changes to *Exclusive-ReadWrite*. This is because no lock is held and C_d is empty. The second part of increment operation "write" is done by the same thread, and doesn't cause any transition (if it is not interfered by the worker thread).

Exclusive-ReadWrite: Location d is accessed by multiple threads and the lockset discipline alone is not sufficient. We track the thread segment identifier corresponding to the most recent access to d. Similar to *Exclusive-Read* or *Exclusive-Write*, we remain in this state as long as there is a hb-relation between successive accesses. When there is a write operation and $TS_d \parallel TS_t$, we enter *Shared-Modified2*. When a read operation happens and there is a hb-relation, we return to *Shared-Read*.

```
State Exclusive-ReadWrite:
  When $t$ executes Read(d):
    if $TS_d \parallel TS_t$ then
      $C_d \leftarrow L_t$
      $S_d \leftarrow \{t,\ threadof(TS_d)\}$
      set state to Shared-Modified2
    end
    else
      set state to Shared-Read
    end
  end

  When $t$ executes Write(d):
    if $TS_d \xrightarrow{hb} TS_t$ then
      $TS_d \leftarrow TS_t$
      set state to Exclusive-ReadWrite
    end
    else if $TS_d \parallel TS_t \wedge L_t \neq \emptyset$ then
      $TS_d \leftarrow TS_t$
      $C_d \leftarrow L_t$
      $S_d \leftarrow \{t,\ threadof(TS_d)\}$
      set state to Shared-Modified2
    end
    else
      set state to Race
    end
  end
```

If the increment operation (line 21) is interfered by the write/read operation of the worker thread (lines 37,38) while being in *Exclusive-ReadWrite*, we enter *Race*, because no hb-relation exists between threads and parallel write accesses happen. If there was some kind of synchronization such as lock/unlock between accesses by threads, we would enter *Shared-Modified2*.

Shared-Modified2: Location d is concurrently read and written by multiple threads. When entering this state, the lockset C_d is initialized to L_t. Both lockset and thread segments are tracked. If the lockset is empty and d is concurrently accessed, generate an error and enter state *Race*. This is the only state where both the happens-before relation and the lockset is analyzed. But whenever a happens-before relation exists between successive accesses, return to *Exclusive-ReadWrite*. This speeds up the algorithm because it reduces the overhead for locksets, especially in long-running programs.

```
State Shared-Modified2:
When $t$ executes read(d) $\vee$ write(d):
    $C_d \leftarrow C_d \cap L_t$
    if $C_d = \emptyset$ then
        if $TS_d \xrightarrow{hb} TS_t$ then
            $TS_d \leftarrow TS_t$
            set state to
            Exclusive-ReadWrite
        end
        else
            set state to Race
        end
    end
    else
        set state to Shared-Modified2
    end
end
```

The variable GLOB has the state *Race* after the execution of line 21 of main thread and line 37 of worker thread. It remains in this state by any further write/read access (the read access at line 38 does not change the state).

Race: A potential race is detected and reported. This state can be reached from *Shared-Modified2* when $\mathrm{C}_d = \{\}$ and $TS_d \parallel TS_t$, which means that d is concurrently accessed by multiple threads without a common lock. Also it is possible to reach *Race* from all exclusive states, in case a write happens concurrently with another access and $L_t = \{\}$. Transitions on empty L_t prevent false negatives in many situations. L_t is the set of locks currently held by a thread during program execution, and tracking it involves hardly any overhead.

One can see that entering *Race* while being in a shared or exclusive state requires at least *two* successive unsynchronized accesses. We will see in the second variant of MSM presented in the following section, only having *one* unsynchronized access is sufficient to enter *Race* resulting in a more sensitive detection algorithm.

5.2.4. Memory State Machine for Short-running Applications

In the previous section, we presented a race detection approach that significantly reduces the false alarms [23]. This approach is suitable for analyzing long-running applications without overwhelming the user with false alarms. In long-running applications, a data race pattern is likely to be repeated. Based on this assumption, the race detector presented before defers certain race reports until the race reoccurs, thus reducing false alarms.

The example provided in Figure 5.3 shows a false positive occurs on variable GLOB. The read operations on GLOB at the end are in parallel, but do not create any data race. MSM-long prevents this false positive and waits for a further unsynchronized write access. By introducing the states *Exclusive-ReadWrite* and *Shared-Modified1* the false positive on GLOB is not reported. See Table 5.2 for the stepwise trace of accesses to GLOB.

```
int GLOB = 0;

int main(){

  create(threadid, worker);

  lock(l);
    GLOB = 1;
  unlock(l);

  sleep (2000);

  printf(GLOB);

  join(threadid, NULL);
}
```

(a) main thread

```
void worker(){

  sleep (2000);

  lock(l);
    GLOB++;
  unlock(l);

  sleep (4000);

  printf(GLOB);

  return NULL;
}
```

(b) worker thread

Figure 5.3.: Preventing a false positive on the shared variable GLOB by MSM-long.

But what happens if the program runs briefly? Then, races may not occur several times. This situation could happen especially during programming by unit testing. For this reason, we introduce a more sensitive race detection algorithm to discover races even if they occur only once. We propose another memory state machine which is more suitable for short-running applications[21]. The new state machine concentrates on accurately detecting data races and

Line No.	GLOB	Old State	New State	$\xrightarrow{hb}$	L_t	C_d
main(): 1	Initialization	-	New	-	-	-
main(): 8	Write	New	Excl-W	-	{}	-
worker(): 22	Read	Excl-W	Sh-M1	-	{l}	{l}
worker(): 22	Write	Sh-M1	Sh-M1	-	{l}	{l}
main(): 13	Read	Sh-M1	Excl-RW	-	{}	{}
worker(): 27	Read	Excl-RW	Sh-M2	no	{}	-

Table 5.2.: MSM-long does not report the false positive on variable GLOB depicted in Figure 5.3.

prevents false negatives while avoiding false positives as demonstrated by the results in Section 9. The new state machine is called *MSM-short* which is depicted in Figure 5.4 and explained in the following section. Compared to MSM-long, MSM-short has two states less.

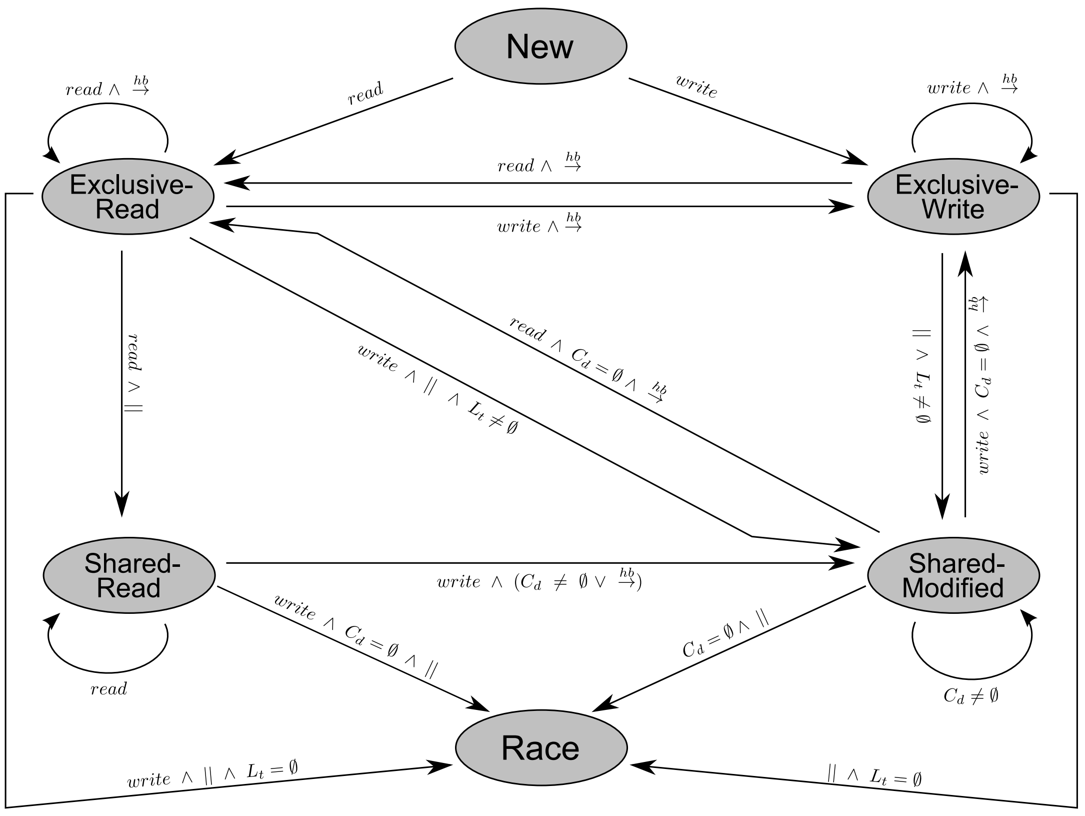

New
$read \wedge \xrightarrow{hb}$
$read$
$write$
$write \wedge \xrightarrow{hb}$
Exclusive-Read
$read \wedge \xrightarrow{hb}$
$write \wedge \xrightarrow{hb}$
Exclusive-Write
$read \wedge ||$
$write \wedge || \wedge L_t \neq \emptyset$
$read \wedge C_d = \emptyset \wedge \xrightarrow{hb}$
$|| \wedge L_t \neq \emptyset$
$write \wedge C_d = \emptyset \wedge \xrightarrow{hb}$
Shared-Read
$write \wedge (C_d \neq \emptyset \vee \xrightarrow{hb})$
Shared-Modified
$read$
$write \wedge C_d = \emptyset \wedge ||$
$C_d = \emptyset \wedge ||$
$C_d \neq \emptyset$
Race
$write \wedge || \wedge L_t = \emptyset$
$|| \wedge L_t = \emptyset$

Figure 5.4.: Memory state machine for short-running applications.

5.2.5. Principles of MSM-short

As Figure 5.4 depicts, the main idea is to avoid entering a shared state until the happens-before analysis shows that there are concurrent accesses to a memory location. Threadset and lockset tracking are performed only in shared states. No thread segment tracking is performed in *Shared-Read*. Only state *Shared-Modified* requires both lockset updates and happens-before analysis. Tracking both lockset and thread segments for each access during program execution can be quite expensive in both time and space (memory consumption).

For this reason, the happens-before analysis in *Shared-Modified* is deferred until the lockset of a location is empty. That is, we do not track the thread segments until the lockset would report a race, leading to performance improvement. If there is a happens-before relation, we return to one of the exclusive states.

As in MSM-long, separate *Exclusive-Read* and *Exclusive-Write* states are beneficial for several reasons. The state machine can distinguish a read after a write or a write after a read. We have more information about the accesses in the past, making the detector work more precisely. In addition, this distinction helps the detector to handle races that could happen only once during initialization time[23]. Missing such races is a shortcoming in current Eraser-style detectors.

5.2.6. States of MSM-short

As in MSM-long, thread segments (e.g. TS_t) are used to indicate a hb-relation between two successive accesses to a memory location. Some states are similar to the states in MSM-long. Here we discuss the differences to the MSM-long algorithm.

A scenario for a data race is depicted in Figure 5.5 which is used for the explanation of different states. The scenario is similar to previous one (Figure 5.5), but it is shorter and a race on the shared variable GLOB happens only after few accesses by threads. Note that the race in this scenario is not detected by MSM-long, as it occurs only once. It would be detected if repeated. But we are able to detect the race by MSM-short.

As shown in Table 5.3, variable GLOB has the state *New* after initialization.

New: Newly allocated memory that is not yet accessed. As in MSM-long, no lockset is needed. On the first *write/read* access, we enter state *Exclusive-Write/Exclusive-Read*..

```
 1 int GLOB = 0;
 2 int COND = 0;
 3
 4 int main(){
 5
 6   create(threadid, worker);
 7
 8   GLOB = 1;
 9   printf(GLOB);
10
11   lock(l);
12   while(COND !=1){
13     wait(cv);
14   }
15   unlock(l);
16
17   GLOB++;
18
19   join(threadid, NULL);
20 }
```

(a) main thread

```
21 void worker(){
22
23   sleep (2000);
24
25   printf(GLOB);
26
27   lock(l);
28     COND++,
29     signal(cv);
30   unlock(l);
31   sleep (4000);
32
33   printf(GLOB);
34
35   return NULL;
36 }
```

(b) worker thread

Figure 5.5.: A data race occurs on the shared variable `GLOB` after only **one** unsynchronized access - overall *seven* accesses.

```
State New:
  When $t$ executes read(d) $\vee$ write(d):
    $TS_d \leftarrow TS_t$
    set state to Exclusive-Read / Exclusive-Write
  end
```

After the first write and read accesses by the main thread (lines 8, 9), `GLOB` has *Exclusive-Write* and *Exclusive-Read* respectively.

Exclusive-Write: Location d is synchronized with hb-relations and the last access was a write by a particular thread segment. No lockset is needed. We remain in exclusive state as long as successive accesses are ordered by hb-relation, since there are no concurrent accesses to d. When a write or read occurs which is parallel to previous access, we enter *Shared-Modified*. Otherwise, we switch to *Exclusive-Write* or *Exclusive-Read* corresponding to the type of current operation.

It is possible to reach *Race* from exclusive states, in case an access happens concurrently with another access and $L_t = \emptyset$ (i.e. an unsynchronized access happens). Transitions from exclusive states to *race* if L_t is empty

Line No.	GLOB	Old State	New State	$\overset{hb}{\rightarrow}$	L_t	C_d
main(): 1	Initialization	-	New	-	-	-
main(): 8	Write	New	Excl-W	-	{}	-
main(): 9	Read	Excl-W	Excl-R	yes	{}	-
worker(): 25	Read	Excl-R	Sh-R	no	{}	{}
main(): 17	Read	Sh-R	Sh-R	-	{}	{}
main(): 17	Write	Sh-R	Sh-Mod	no	{}	{}
worker(): 33	Read	Sh-Mod	Race	no	{}	{}

Table 5.3.: Catching a data race on variable GLOB after *one* unsynchronized access by MSM-short (depicted in Figure 5.5).

prevent false negatives in many situations. If $L_t = \emptyset$ were not checked, we would have false negatives. L_t is the set of locks currently held by a thread during program execution, and tracking it involves hardly any overhead.

```
State Exclusive-Write:
When $t$ executes read(d) $\vee$ write(d):
    if $TS_d \overset{hb}{\rightarrow} TS_t$ then
        $TS_d \leftarrow TS_t$
        set state to Exclusive-Read / Exclusive-Write
    end
    else if $TS_d \mid\mid TS_t \wedge L_t \neq \emptyset$ then
        $TS_d \leftarrow TS_t$
        $C_d \leftarrow L_t$
        $S_d \leftarrow \{t,\ threadof(TS_d)\}$
        set state to Shared-Modified
    end
    else
        set state to Race
    end
end
```

Exclusive-Read: Similar to *Exclusive-Write*, except that the last access was a read operation. Location d is synchronized with hb-relations and no lockset is needed.

When a parallel access occurs, this is potentially a race except in the following cases:

- On a read operation, we enter *Shared-Read*, because parallel reads are not considered a data race. At this moment the happens-before chain is broken. Thread segment TS_d is kept to be used in *Shared-Read.*

- On a write operation, if the thread holds a lock, we assume that from now on, variable d is protected by locks and we enter *Shared-Modified.*

In all other cases, we report a race.

State Exclusive-Read:

```
When $t$ executes read(d):
    if $TS_d \parallel TS_t$ then
        keep $TS_d$
        $C_d \leftarrow L_t$
        $S_d \leftarrow \{t,\ threadof(TS_d)\}$
        set state to Shared-Read
    end
    else
        set state to Exclusive-Read
    end
end
```

```
When $t$ executes write(d):
    if $TS_d \overset{hb}{\rightarrow} TS_t$ then
        $TS_d \leftarrow TS_t$
        set state to Exclusive-Write
    end
    else if $TS_d \parallel TS_t \wedge L_t \neq \emptyset$ then
        $TS_d \leftarrow TS_t$
        $C_d \leftarrow L_t$
        $S_d \leftarrow \{t,\ threadof(TS_d)\}$
        set state to Shared-Modified
    end
    else
        set state to Race
    end
end
```

By execution of the parallel read access to GLOB by the worker thread at line 25, we enter *Shared-Read.*

Shared-Read: Location d is concurrently accessed by multiple threads, but all accesses are reads. *Shared-Read* allows parallel reads. We enter this state from *Exclusive-Read* when a read results in multiple concurrent accesses.

In this state, thread segments are not updated. We track only the lockset C_d, which is initialized to L_t and the threadset to see if a variable is shared between threads. The lockset is updated for every access. If a write occurs, enter *Shared-Modified*, except when this write is in parallel with the TS_d stored in *Exclusive-Read*, and no lock protects it.

This situation is the only case of reporting a race in *Shared-Read.* Since we do not update the thread segment in *Shared-Read*, the thread segment in *Exclusive-Read* is stored at the point where the happens-before chain is broken. When a write causes a transition, the thread segment of the writing thread and the stored TS_d are examined. If there are parallel accesses and the lockset is empty, we enter *Race.* This increases the chance of detecting races raised in *Shared-Read*, unlike many Eraser-style tools that lack the ability to detect races for shared-read data.

State Shared-Read:

```
When $t$ executes read(d):
    $C_d \leftarrow C_d \cap L_t$
    $S_d \leftarrow S_d \cup \{t\}$
    set state to Shared-Read
end

When $t$ executes write(d):
    if $TS_d \overset{hb}{\rightarrow} TS_t$ then
        $C_d \leftarrow C_d \cap L_t$
        $S_d \leftarrow S_d \cup \{t\}$
        set state to Shared-Modified
    end
    else
        set state to Race
    end
end
```

The increment operation executed on GLOB at line 17 causes a transition to *Shared-modified.* This is because no lock is held by the main thread during incrementing. Also, no hb-relation exists between the thread segment of the current write operation and the last stored thread segment.

Shared-Modified: Location d is concurrently read and written by multiple threads. We presume that variable d is protected by the locks in C_d. If it is entered from an exclusive state, the lockset C_d is initialized to L_t. If this state is entered from *Shared-Read*, the lockset is taken from this state. In addition to threadset, both lockset and thread segments are tracked.

If the lockset C_d is empty, then d is not correctly synchronized with locks. So we check if other synchronization patterns impose any happens-before relations. If not, we generate an error and enter the *Race* state. If there are hb-relations, we return to an exclusive state. This speeds up the algorithm because it reduces the overhead for lockset analysis for all accesses during program execution.

Note that this is the only state where both the happens-before relations and the locksets are analyzed.

```
State Shared-Modified:
    When $t$ executes read(d) $\vee$ write(d):
        $C_d \leftarrow C_d \cap L_t$
        if $C_d = \emptyset$ then
            if $TS_d \overset{hb}{\rightarrow} TS_t$ then
                $TS_d \leftarrow TS_t$
                set state to Exclusive-Read / Exclusive-Write
            end
            else
                set state to Race
            end
        end
        else
            set state to Shared-Modified
        end
    end
```

The last read access done by the worker thread (line 33) causes a data race. This is caught by MSM-short and we enter *Race*.

Race: A potential race is detected and reported. Introducing this separate state is useful, because once the race is reported, the tool does not spend time on this memory location any more.

The detected race on `GLOB` is not detected by MSM-long, since no race is reported in *Shared-modified1* by MSM-long. The race would be reported by MSM-long, if it repeated or further unsynchronized read/write accesses happen.

5.2.7. Discussion and Comparison of Memory State Machines

MSM-long has two different *Shared-Modified* states, whereas MSM-short has only one; therefore, it is more sensitive [21]. The two distinct states in MSM-long were introduced to defer race warnings. It is assumed that in long-running applications, races on a memory location happen several times. So, on first occurrence of a potential race, MSM-long waits until it is repeated to be sure of reporting real races.

In contrast, MSM-short will warn about races immediately after the first indication of incorrect synchronization. That is, there is a difference in the way the empty lockset is handled. In *Shared-Modified1* in MSM-long, a transition to *Race* on an empty lockset would lead to numerous false positives (as shown in the example provided in Figure 5.3). On the other hand, if the alarm produced in the single *Shared-Modified* in MSM-short is a true positive, it could be caught immediately avoiding a possible false negative.

We examined both versions on various benchmarks and chose the best solution. Based on our experimental results, the distinction between *Shared-Modified1* and *Shared-Modified2* is beneficial for long-running applications. If an empty lockset in *Shared-modifed1* is indeed a true positive, it is mostly followed by another parallel read/write leading to the *Race* state. MSM-short is also suitable for unit testing and during development when a race pattern may occur only once, and MSM-long is better suited for integration testing.

The distinction between *Shared-Modified1* and *Shared Modified2* in MSM-long is also justified by performance reasons. In *Shared-Modified1*, only the lockset is maintained, whereas in *Shared-Modified2* both lockset and thread segments (happens-before analysis) are tracked. This optimization makes the detector practical for real-world applications.

5.3. Limitations

Our memory state models are a compromise between detecting all races and reporting few false positives. There are some special cases in which a false negative could occur:

One scenario is given in Figure 5.6, when a variable `GLOB` in *New* state is not protected and is initialized by a thread. There is a race condition when a second thread writes to `GLOB` concurrently. If the second thread holds an unrelated lock, a race is caused which will be missed. However, any unprotected read or write after this operation will be reported as a data race. Usually, programmers tend to not use locks for initializations. Thus, the race could be counted as an intentional race.

Note that by correct detection of synchronization with inter-thread event notifications finding some special race patterns are possible. We are able to find races caused by spurious wake ups and further reduce the false negatives (more details are presented in next chapter).

There still are some false positives, and reducing the false positive rate even further is a challenge. This is mainly due to the missing information about synchronization operations (e.g. ad-hoc synchronizations) in the program by the detector that causes an imprecise happens-before detection.

```
int GLOB = 0;

int main()
{
  create(threadid, worker);
  GLOB = 0;
  join(threadid, NULL);

  return 0;
}
```

(a) parent thread

```
void worker()
{
  lock(l);
    GLOB = 1;
  unlock(l);

  return NULL;
}
```

(b) worker thread

Figure 5.6.: Example for a false negative due to the limitation of MSMs.

5.3.1. Imprecise Happens-Before Detection

Reconstructing hb-relations is an important task for race detectors. If hb-relations between thread segments are missed, the detector considers them to be parallel although they were correctly synchronized, thus producing false positives. On the other hand, if we add false relations to the happens-before graph, the detector causes false negatives.

The knowledge of all synchronization operations is crucial to data race detection. We discovered three major problems which lead to inaccuracies in the happens-before graph during data race detection:

- Inter-thread event notifications
- Ad-hoc synchronizations
- Unknown synchronization primitives from unsupported libraries

The problems occur often in parallel programs. The first problem is caused by using inter-thread event notifications via condition variables. The second one deals with ad-hoc synchronization constructs defined by the programmer in program code. The third issue is caused by using unknown synchronization primitives from other libraries that are not supported by the detector.

We discuss these problems in detail in the following chapters and present our solution, which establishes hb-relations and automatically extracts the correct synchronizations induced without relying on annotations or user assistance. Our solution eliminates almost all cases of false alarms and missed races related to inter-thread event notifications, ad-hoc synchronizations and unknown synchronization primitives of unsupported libraries.

Chapter 6.

Detecting Inter-thread Event Notifications

When programmers use inter-thread event notifications, they use two separate parts for synchronization: Functions for event notifications between threads which is provided by a library (e.g. **signal**/**wait** primitives), and condition variables that are implemented in program code. The programmer uses both of them to ensure reliable synchronization. The resulting synchronization is not only dependent on library function calls, but also on the condition variables implemented in the code. Due to that, detectors are not aware of condition variables and have to be cautious when handling inter-thread event notifications. Synchronization with inter-thread event notifications reveals two problems during happens-before analysis: *lost signals* and *spurious wake ups*. Lost signals happen when the signal is sent before a thread waits for it. Spurious wake ups result from the usage of only one condition variable to signal several different conditions at the same time.

By merely intercepting **signal**/**wait**, it is not clear if the synchronization was really successful. Therefore, detectors cannot establish the required hb-relations. This problem does not happen with other types of synchronization primitives. For example, when using locks, the whole synchronization takes place within the library primitive calls **lock** and **unlock**; after the termination of these primitives, it is clear that the synchronization is successfully terminated.

The problems above were not solved in other approaches, e.g. [36], that extend data race detection for condition variables.

6.1. Motivating Example

A higher sensitivity usually also means a higher rate of false alarms. To avoid this, the detector has to distinguish more accurately between real data races and harmless accesses. Only parallel accesses can lead to data races, so the detector has to find out how accesses are ordered.

In some cases, it is extremely difficult to reconstruct the implicit ordering imposed by synchronization primitives, as shown in Figure 6.1. Thread 1 operates on DATA and then signals Thread 2 to take over the data for further processing. The threads are properly synchronized, but there is an ordering in which the happens-before relation involving `signal()` and `wait()` is not visible to the race detector, causing it to issue a false warning on DATA. This situation is as follows.

```
DATA++

lock(l)
  FLAG = 1
  signal(CV)
unlock(l)
```

(a) Thread 1

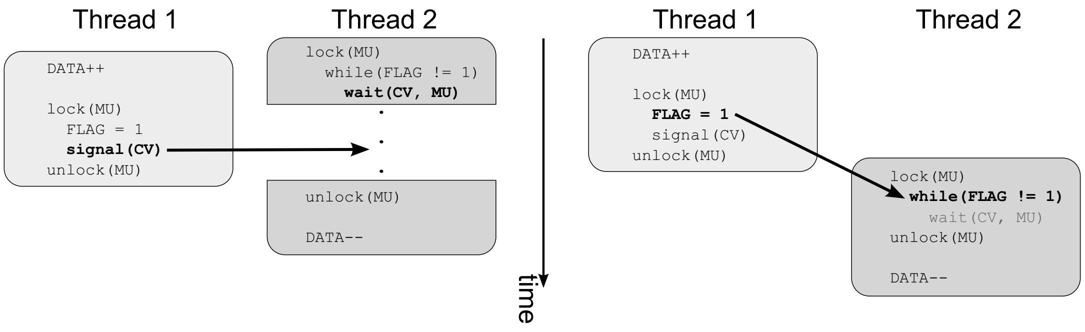
```
lock(l)
  while(FLAG != 1)
    wait(CV)
unlock(l)

DATA--
```

(b) Thread 2

Figure 6.1.: Lost signal when using synchronization primitives `signal()` and `wait()`.

If Thread 1 finishes first and then Thread 2 starts, Thread 2 will not call `wait()`. Consequently, the signal sent by Thread 1 is lost. The instrumentation of `signal()` and `wait()` thus does not help the detector to identify the proper ordering of the threads, since a part of synchronization (`wait()` function) for building the hb-relation is missing. Thread 2 carries on and as soon as it accesses DATA, a data race is reported, even though there is none.

The proper ordering is enforced by the condition variable `FLAG`, but noticed by neither lockset nor happens-before detectors. This scenario is depicted on the right side of Figure 6.2. Another Scenario shown on the left side of Figure 6.2 shows that Thread 2 is executed earlier and is waiting for the signal sent by Thread 1. In this case the functions `signal()` and `wait()` are explicitly called and could be intercepted by the detector to construct the synchronization.



```
Thread 1                      Thread 2
DATA++                        lock(MU)
                                while(FLAG != 1)
lock(MU)                          wait(CV, MU)
  FLAG = 1                          .
  signal(CV) ---------------->      .
unlock(MU)                          .
                              unlock(MU)

                              DATA--
```

```
Thread 1                      Thread 2
DATA++

lock(MU)
  FLAG = 1 ----------------.
  signal(CV)               |
unlock(MU)                 v
                              lock(MU)
                                while(FLAG != 1)
                                  wait(CV, MU)
                              unlock(MU)

                              DATA--
```

Figure 6.2.: Depending on the interleaving during execution, different parts of the code are responsible for constructing the synchronization.

6.2. Spurious Wake ups

Figure 6.3 shows another problematic situation with two pairs of signaling and waiting threads (S_1,W_1) and (S_2,W_2). Each pair accesses a different set of data and uses different conditions. But only one condition variable `CV` is used for conceptually two different signals. This way, a signal would wake up both pairs, which means one signal would spuriously wake up a wrong thread. The situation is also the same if the synchronization primitive **broadcast** is applied instead of **signal**.

```
X++

lock(l1)
  FLAG1 = 1
  signal(CV)
unlock(l1)
```

(a) S_1

```
lock(l1)
  while(FLAG1!=1)
    wait(CV)
unlock(l1)

X--
```

(b) W_1

```
Y++

lock(l2)
  FLAG2 = 1
  signal(CV)
unlock(l2)
```

(c) S_2

```
lock(l2)
  while(FLAG2!=1)
    wait(CV)
unlock(l2)
X--
Y--
```

(d) W_2

Figure 6.3.: Several signaling and waiting threads using the same condition variable `CV`.

Based on the program semantics, (S_1, W_1) are synchronized through variable

`FLAG1` and (S_2, W_2) through `FLAG2`. Thus, after synchronizations W_1 is allowed to access only `X` and W_2 only `Y` . However, W_2 also accesses `X`, and a data race takes place between the pair S_1 and W_2.

The above data race cannot always be identified by the detector, since in some interleavings the detector deduces by mistake that the pair (S_1, W_2) are synchronized. Assume the scenario that W_1 and W_2 are executed and waiting for signals. Afterwards, S_1 finishes and sends a signal that wakes up both threads W_1 and W_2; the detector cannot decide which waiting threads should be synchronized with S_1, and synchronizes both of them with S_1. Since both waiting threads W_1 and W_2 use the same `CV`, it is not possible for the detector to identify the proper ordering, only by intercepting the **wait**`(CV)` function call. It cannot distinguish the correct synchronization and determine if a real synchronization happens.

In the following sections, we show how to solve these problems without relying on source code annotations.

6.3. Detecting Lost Signals

Detecting hb-relations based on intercepting library function calls works fine with the synchronization primitives such as fork/join and barriers. They are explicitly called and could be intercepted directly by a detector. This direct interception is not always the case when using inter-thread event notifications (signal/wait functions). Because the exact point of time that the synchronization is successfully terminated by them is not clear for the detector. This point is determined by the "condition variables" defined by the programmer in the source code and not by the synchronization primitives i.e. `signal()` and `wait()`.

Thus, just by termination of synchronization primitives, a part of the semantics of synchronization is hidden for the race detector. In Section 6.1, we stated a typical synchronization pattern using condition variables. We discussed the problem that condition variables are stateless and signals are not accumulated. If the waiting thread is scheduled after the signaling thread, the signal will be *lost*; the waiting thread will never call the corresponding `wait()`. Thus, the detector *does not know* that synchronization has taken place.

Observing synchronization via condition variables, you will find a similar pattern in most cases that could be easily found in source code: a while-loop that checks a condition and executes the `wait()` function until the while-condition changes. Therefore, the point when the synchronization is successfully terminated is directly after the while loop.

A method may use source code annotations to mark the synchronization point after the while loop. However, this method is troublesome for the developer who uses the detector. The source code of the program has to be available and it has to be changed. Furthermore, the additional compilation process makes this method intrusive in the build process.

In this work, we propose a new automatic method that finds the synchronization points through binary instrumentation. The method could be integrated directly into a dynamic race detector and doesn't need any user intervention. Our method searches for *while-loops* in the binary code which contain a call to the `wait()` library function. All the `wait()`-enclosing loops are instrumented so that the detector will be aware of them. Thus, it does not matter anymore whether a call to `wait()` is executed or not. The detector is able to intercept all `wait()` function calls, even if the signal is sent before calling the `wait()` function. So, it is able to construct the hb-relation based on the corresponding intercepted `signal()/wait()` primitives and lost signals do not affect the detector.

The method works well since the code to correctly wait for a signal (as shown in Figure 6.1) follows specific rules:

- The `wait()` library function is called within a loop, when a specific condition is met.
- It is assured that the signal was sent at the time the program leaves the loop.
- Locks are used to protect the condition variables.

6.3.1. Source Code Annotation

As discussed earlier, a workaround to address the problem of lost signals is source code annotation. An automatic method for source code annotation is described in [30]. If condition variables are used in the program, the method marks synchronization points in source code. First, it searches for a sequence of **`lock`** ... **`while`** ... **`wait`**`(cv)` ... **`unlock`** in the source code. When it finds one, it annotates the end of the while loop by inserting `ANNOTATE_WAIT(cv)`. This annotation gives the detector the necessary information for the synchronization, without affecting the program semantics. Figure 6.4 displays a simple example of source code annotation.

```
lock (mu)
...
while(condition)
{
    ...
    wait(cv, mu)
    ...
}
ANNOTATE_WAIT(cv)
...
unlock(mu)
```

Figure 6.4.: Annotated while loop when using condition variable.

The detector handles the `ANNOTATE_WAIT(cv)` in the same way as a call to the `wait(cv)` function. At the end of the call, the hb-relation between the synchronization point within the signaling thread and the waiting thread is constructed. Compared to the `wait(cv)` function inside the while loop, the function `ANNOTATE_WAIT(cv)` will be called each time, so the hb-relation is guaranteed to exist.

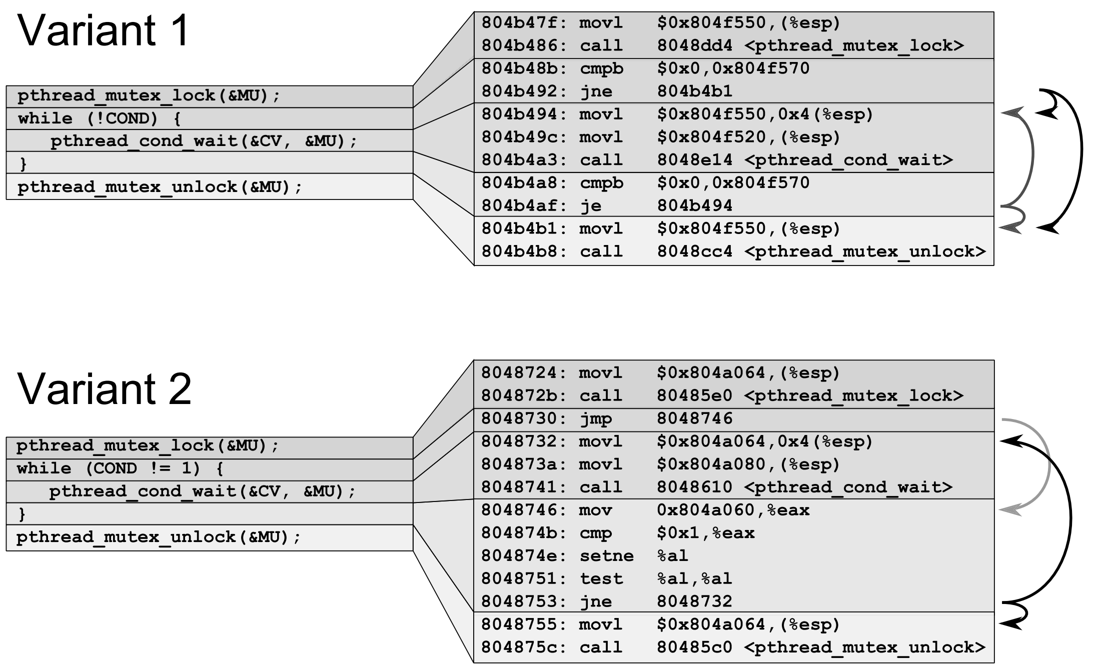
6.3.2. Binary Instrumentation

Locating the synchronization pattern (`lock` ... `while` ... `wait(cv)` ... `unlock`) described above in binary code is not trivial. In binary level, the high level source code is compiled and transformed into a set of machine instructions. Constructing a proper hb-relation at this level requires having the correct synchronization points.

Finding Synchronization Points

In binary instrumentation, the end of the while-loops have to be found to be considered as synchronization points. However, in binary level one cannot find while-loops, since after the build process they are converted to *branches* or *jump* instructions.

As an example, Figure 6.5 depicts two different variants of while-loops in machine code. A while-loop is transformed to one or several machine code blocks. There is a code block containing a conditional branch (jump). Depending on the branch condition the destination code block is chosen. In Variant 2, one can find only one conditional block at the end of the code, whereas in Variant



```
pthread_mutex_lock(&MU);
while (!COND) {
   pthread_cond_wait(&CV, &MU);
}
pthread_mutex_unlock(&MU);
```

```
804b47f: movl   $0x804f550,(%esp)
804b486: call   8048dd4 <pthread_mutex_lock>
804b48b: cmpb   $0x0,0x804f570
804b492: jne    804b4b1
804b494: movl   $0x804f550,0x4(%esp)
804b49c: movl   $0x804f520,(%esp)
804b4a3: call   8048e14 <pthread_cond_wait>
804b4a8: cmpb   $0x0,0x804f570
804b4af: je     804b494
804b4b1: movl   $0x804f550,(%esp)
804b4b8: call   8048cc4 <pthread_mutex_unlock>
```

```
pthread_mutex_lock(&MU);
while (COND != 1) {
   pthread_cond_wait(&CV, &MU);
}
pthread_mutex_unlock(&MU);
```

```
8048724: movl   $0x804a064,(%esp)
804872b: call   80485e0 <pthread_mutex_lock>
8048730: jmp    8048746
8048732: movl   $0x804a064,0x4(%esp)
804873a: movl   $0x804a080,(%esp)
8048741: call   8048610 <pthread_cond_wait>
8048746: mov    0x804a060,%eax
804874b: cmp    $0x1,%eax
804874e: setne  %al
8048751: test   %al,%al
8048753: jne    8048732
8048755: movl   $0x804a064,(%esp)
804875c: call   80485c0 <pthread_mutex_unlock>
```

Figure 6.5.: While-loops in machine code.

1, there are two different blocks at the top and bottom of the code blocks containing the conditional branches. In Variant 2, the compiler has optimized the code and duplicated the conditional branches.

After the evaluation of conditional branch, either the loop-body will be executed or the loop is jumped over and terminated. Thus, the program counter after the execution of the branch could be set to two possible instruction sets: Either to the beginning address of the loop body or the ending address of the whole loop. If one instruments the instruction at the ending address of the loop body, the race detector is notified of the end of the loop. I.e. the detector will be aware of termination of a successful synchronization via condition variable at runtime. To assure identifying synchronization with condition variables, all conditional branches in the program have to be examined. Additionally, these branches are checked to see if they enclose any **wait**(cv) functions, since only loops applied for synchronization with condition variables have to be considered.

Determining the Parameters

So far, our method can detect *where* a signal should be waited for. But it still does not know exactly, *which* signal to wait for. In fact, after finding the synchronization point of the waiting side, some information is still missing, namely which signaling side belongs to this synchronization point. This information is given by the parameters provided in the **wait**`(cv)` function, i.e. the condition variable used in the synchronization function (in this case, only the parameter *cv* is used). Therefore, we wait for the **signal**`(cv)` which used the same parameters as in the synchronization function **wait**`(cv)`. In this way, we are able to specify both synchronization points related to condition variable *cv*.

The parameters are crucial to set up the hb-relation correctly. A possible way to get these parameters is to wait for a call to **wait**`(cv)` and store its parameters. This causes problems, because the **wait**`(cv)` function will not be executed each time as in the cases with lost signals.

A solution to overcome these problems is on-the-fly simulation of the loop body execution, and do a stack walk to extract the parameter set of **wait**`(cv)` function. The simulation has to be done without side effects. The method works, except for a few implementation limitation that we will discuss in Chapter 8.

At the end we have combined both methods to guarantee reliable identification of condition variables during the instrumentation. Firstly, the beginning of the while-loop will be simulated to get the set of parameters. Upon termination of the while-loop, if the **wait**`(cv)` function is executed the parameters provided by the **wait**`(cv)` function are used. Otherwise, if the **wait**`(cv)` function is not executed at all (i.e. lost signal cases), the parameter set of **wait**`(cv)` function derived from the simulation is used.

6.4. Data Dependency

In Figure 6.3, we presented an example in which the condition variable `cv` was used by different threads: Two pairs of signaling and waiting threads accessed different data, but only one condition variable was used. That is, the parts of synchronization done in source code use only one `cv` for both pairs. The synchronization between four threads depicted in Figure 6.3 builds the hb-relations demonstrated in Figure 6.6(a). There is no path between the thread groups (S_1, W_1) and (S_2, W_2). Therefore, the pair (S_1, W_1) is parallel with the pair (S_2, W_2), but ($S_1 \overset{hb}{\rightarrow} W_1$) and ($S_2 \overset{hb}{\rightarrow} W_2$).

In fact, race detectors store only one synchronization point for each conditional variable. Therefore, when sending several signals in sequence, the older point will be overwritten each time. For instance, assume threads are executed as in the following sequence: S_1, S_2, W_2, W_2. The detector constructs a false hb-relation as shown in Figure 6.6(b), since the synchronization point of S_1 is overwritten by S_2. In this case, a false warning on `X` between S_1 and W_1 is reported due to the false happens-before graph.

An elementary solution is to combine all synchronization points with artificial synchronization points [30]. For this reason, we could introduce *fake thread segments*. By introducing a fake segment for `cv`, happens-before edges are redirected to and from a fake segment as in Figure Figure 6.6(c): The calculated happens-before graph contains real hb-relations. Still some additional false hb-relations are constructed, which are not caused by any synchronization construct in the program. For instance, the false hb-relation between S_1 and W_2 masks the race on harmful access on `X` by W_2 and causes a false negative.

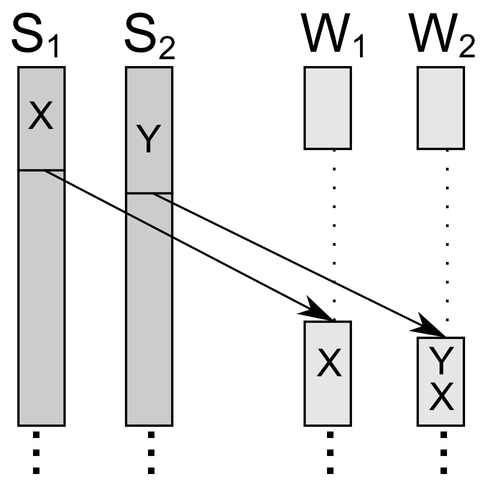
To identify the correct synchronization, detectors must be able to understand the semantics of conditions implemented in the program code. This could be estimated by a data dependency analysis between the signaling and waiting threads. If a thread, shortly before calling a **signal** or **broadcast**, changes a variable and after that the same variable is checked within a condition by another thread, apparently there is a causal relation between both threads. We call this relation a *write/read relation* or shortly *wr-relation*. In compiler terminology this is called *Definition-Use chain*. A Definition-Use chain (DU-chain) consists of a definition, D, of a variable and all the uses, U, reachable from that definition without any other intervening definitions.

6.4.1. WR-Relation

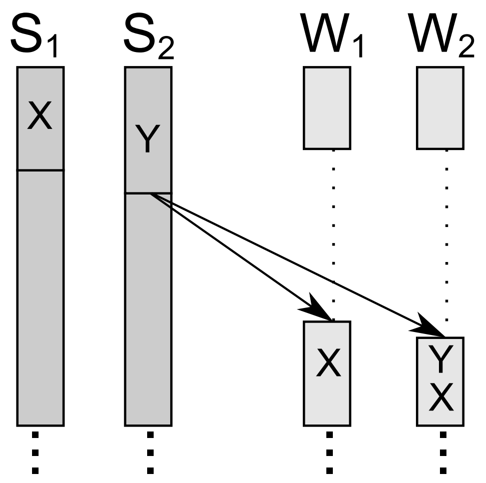
In our sample, there are only wr-relations within the pairs (S_1, W_1) and (S_2, W_2). The signaling threads modify their corresponding `FLAG` variable and the waiting threads evaluate them in their conditions. If we indicate the hb-relation between threads based on the existing data dependency, the only correct happens-before graph could be achieved easily. Figure 6.6(d) depicts the correct hb-relations that Helgrind$^+$ identifies [21]. No false positives or false negatives are caused by this graph. Note that the synchronization points within the fake thread segment do not create any data dependency.

Identifying wr-relations at runtime requires monitoring all variables adjacent to a **signal** or a **broadcast**. These variables are protected by a **lock** operation. We need to record only variables between a **lock** operation and the **signal** or **broadcast**. All *write accesses* to variables within this code block will be recorded.

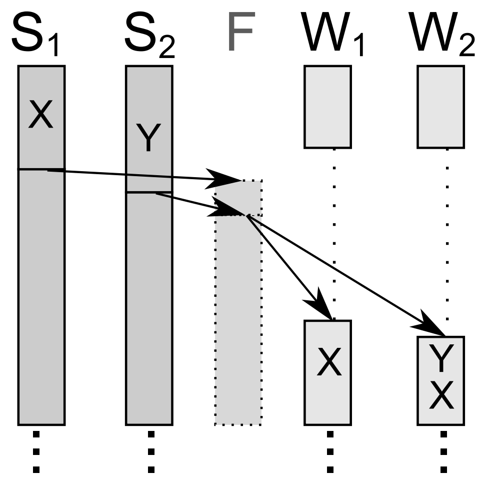
On the side of waiting thread, the conditional jump (branch) is already instrumented. The evaluation of the jump condition is completed at this point. All *read accesses* on variables between the **lock** operation and the *conditional jump* are recorded. When synchronization terminates successfully, we can search for



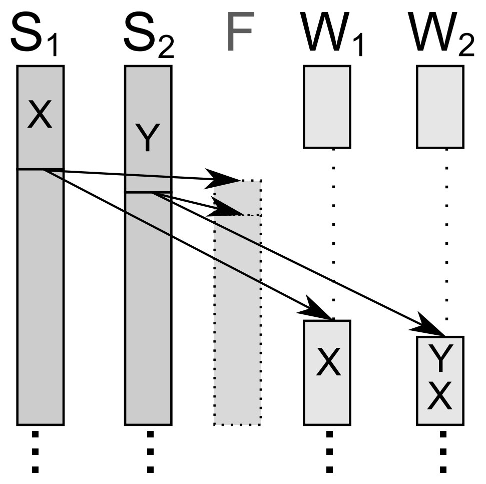
(a) Correct happens-before graph



(b) Only one synchronization point for a condition variable on the signaling side is stored



(c) Introducing fake thread segments (F) that aggregate synchronization points



(d) hb-relations calculated by a data dependency analysis (wr-relation) in Helgrind$^+$

Figure 6.6.: Happens-before graphs generated by different methods to variables X and Y.

direct data dependencies (wr-relations) between write and read accesses. Because, multiple wr-relations could be found, a hb-relation is constructed for each wr-relation.

If no data dependency is found, the fake thread segment is used as an alternative solution, which merges all sent signals of a condition variable into a single synchronization point.

Searching for data dependencies is allowed after finishing a successful synchronization. Under other circumstances, it is not clear whether the signaling thread has already modified the variable or not. Furthermore, it must not have side effects on the normal behavior and the execution of the **wait** function.

Algorithm 2 depicts the extension of our race detection algorithm to identify hb-relations based on *wr-relations*. By executing a lock operation, thread t enters a lock-protected region with **lock()**. We start to record all the following memory accesses. The read memory accesses on variables (such as d) are recorded into a set of read variables called R_t and the write memory accesses are recorded into the set of written variables W_t. R_t and W_t are reset when we leave a lock-protected region with **unlock()**.

```
{ Lock operations }

After t executes lock(l):
    $R_t \leftarrow \emptyset$
    $W_t \leftarrow \emptyset$
end

After t executes unlock(l):
    $R_t \leftarrow \emptyset$
    $W_t \leftarrow \emptyset$
end

{ Memory accesses }

After t executes read(d):
    $R_t \leftarrow R_t \cup \{d\}$
end

After t executes write(d):
    $W_t \leftarrow W_t \cup \{d\}$
end

{ Condition variables }

Before t executes signal(cv):
    foreach $d$ in $W_t$ do
        $DS(d) \leftarrow TS_t$
    end
    $TS_{cv} \leftarrow NewSegment(TS_t, TS_{cv})$
    $TS_t \leftarrow NewSegment(TS_t)$
end

After t executes wait(cv):
    foreach $d$ in $R_t$ do
        if $DS(d)$ exists then
            $TS_t \leftarrow$
            $NewSegment(TS_t, DS(d))$
        end
        else
            $TS_t \leftarrow$
            $NewSegment(TS_t, TS_{cv})$
        end
    end
end
```

Algorithm 2: Basic detection algorithm and lockset update rules with wr-relation.

A call to signal in a lock-protected region causes the detector to scan all the modified variables from the beginning of the lock-region up to signal. It maps all variables recorded in W_t to the current thread segment identifier. This is done by $DS(d)$ in the algorithm, i.e. $DS : d \mapsto ts$ where ts denotes the thread segment of the last signaler which modified the variable d. Additionally, a fake thread segment (TS_{cv})is created which is ordered after every signaler segment. TS_t is the thread segment of the current thread t.

When a wait (or an annotated wait in a loop) is called, all recorded read operations R_t are checked to see if there is a mapping $DS(d)$ indicating any wr-relation on d. If it exists, a hb-relation between the waiting thread and the real signaler is constructed. Otherwise, we have to fall back to use the former technique with the fake thread segment (TS_{cv}), which was created by our signal-handler.

6.5. Summing-up

This section describes the steps for the reliable identification of inter-thread event notifications. The approach is conceptually divided into three phases:

1. *Pre-instrumentation phase*: The relevant locations of machine code are blockwise analyzed and particular points are marked to be instrumented.
2. *Instrumentation phase*: Instrumentation takes place just-in-time and blockwise. The analysis code (analysis functions) is inserted into the machine code.
3. *Runtime phase*: During *runtime* the analysis functions perform the actual program analysis.

6.5.1. Pre-instrumentation Phase

In the pre-instrumentation phase, we look into the given machine code and find all *conditional jumps*. Then, *for each* conditional jump, we specify two possible instruction addresses proceeding after the jump instruction. Regarding the control flow graph, one target is the address of the first instruction inside the loop body and the other one is the instruction address outside the loop body, right after ending the loop. The loop body is located between these two instruction addresses. For this reason, we search for a wait-function call between these addresses. If there is a wait call, we assume that the loop is used for inter-thread event notification via condition variables. We interpret the smaller address as the starting address and the larger one as the ending address of the loop body. The ending address indicates the end of a synchronization and is marked for instrumentation. If no **wait** call is found the conditional jump is ignored. The following list depicts briefly how the code locations are marked:

```
For all conditional jumps in program code:
  body_start ← min(jump_destination1, jump_destination2)
  body_end ← max(jump_destination1, jump_destination2)

  if ∃ wait-function call between body_start and body_end:
    mark body_start
    mark body_end
    mark jump
```

6.5.2. Instrumentation phase

In the previous phase, we specified the conditional jumps with the beginning and the ending address of the whole synchronization. We instrument the program code based on the commands shown in Table 6.1.

Instruction / Markings	Instrumentation before instruction	Instrumentation after instruction
lock		BEGIN_RECORDING_READS BEGIN_RECORDING_WRITES
unlock	STOP_RECORDING_READS STOP_RECORDING_WRITES	
marked conditional jump	STOP_RECORDING_READS GET_PARAMETERS	
marked body_start	START_IGNORING_WAITS	
marked body_end	STOP_IGNORING_WAITS ANNOTATE_WAIT	
signal or **broadcast**	STOP_RECORDING_WRITES	HANDLE_SIGNAL
wait		HANDLE_WAIT

Table 6.1.: Instrumentation commands for the reliable handling of inter-thread event notifications (condition variables).

The commands are functions that are inserted into machine code and will be executed at runtime. The only exception is the command GET_PARAMETERS that executes a sequence of operations to specify the parameters for **wait**-function calls, required by ANNOTATE_WAIT. Details about specifying the parameters of

wait function will be explained in Chapter 8. The column *Instrumentation before instruction* in Table 6.1 lists all commands that are injected before instruction; *Instrumentation after instruction* describes commands injected after the instruction.

Listing 6.1 lists the result of the pre-instrumentation and the instrumentation phase for both loop variants depicted in Figure 6.5.

6.5.3. Runtime phase

The actual program analysis happens at runtime. We record variable accesses at particular regions in the program code to find data dependencies and extract wr-relations.

Functions BEGIN_RECORDING_READS and STOP_RECORDING_READS record read accesses. Only read accesses of the current execution of a lock-protected region are recorded. Thus, BEGIN_RECORDING_READS overwrites the previous recording. Similarly, BEGIN_RECORDING_WRITES and STOP_RECORDING_WRITES record write accesses. Listing 6.2 demonstrates what exactly the injected functions do at runtime.

Listing 6.3 demonstrates the complete instrumented code of a signaling thread. For better understanding, the instrumented machine code is expressed in C/C++ source code. The thread modifies variables that affect the loop condition within a lock protected region. During instrumentation, the region that begins with a **lock** and ends with **signal**, **broadcast** or **unlock** is marked. All write accesses within this region are recorded at runtime. As soon as **signal** or **broadcast** is called, all recorded accesses are associated with the current point of time (related thread segment identifier).

Basically, we have to record only write accesses on signaler side and only read accesses on waiting side. However, as shown in Listing 6.3, after the **lock**-operation all read and write accesses are recorded. This is because at the beginning of a lock-protected region, it is not clear whether later on a **signal** or **wait** will be called.

Listing 6.4 depicts the instrumented code of a waiting thread. In the code block between the **lock** operation and the *conditional jump* (while-loop), variables upon which the condition is dependent are accessed. This code block is also instrumented so that all read accesses will be recorded at runtime. At the end of synchronization, which is marked with ANNOTATE_WAIT, the detector searches for possible data dependencies based on recorded read/write accesses to extract any wr-relation between the signaling and waiting threads.

```
Variant 1:
 804b47f: movl    0x804f550,(%esp)       # [mutex]
 804b486: call   8048dd4                 # pthread_mutex_lock(&mutex
     )
          BEGIN_RECORDING_READS
          BEGIN_RECORDING_WRITES
 804b48b: cmpb    0x0,0x804f570          # [open]
 804b492: STOP_RECORDING_READS
          GET_PARAMETERS
          jne    804b4b1                 # while(!open) {
 804b494: START_IGNORING_WAITS
          movl    0x804f550,0x4(%esp)    # [cond]
 804b49c: movl    0x804f520,(%esp)       # [mutex]
 804b4a3: call   8048e14                 # pthread_cond_wait(&cond,
     &mutex)
          HANDLE_WAIT
 804b4a8: cmpb    0x0,0x804f570          #
 804b4af: je     804b494                 # }
 804b4b1: STOP_IGNORING_WAITS
          ANNOTATE_WAIT
          movl    0x804f550,(%esp)       # [mutex]
 804b4b8: STOP_RECORDING_READS
          STOP_RECORDING_WRITES
          call   8048cc4                 # pthread_mutex_unlock (&
              mutex)


Variant 2:
 8048724: movl    0x804a064,(%esp)       # [MU]
 804872b: call   80485e0                 # pthread_mutex_lock(&MU)
          BEGIN_RECORDING_READS
          BEGIN_RECORDING_WRITES
 8048730: jmp    8048746                 # while(COND != 1) {
 8048732: START_IGNORING_WAITS
          movl    0x804a064,0x4(%esp)    # [CV]
 804873a: movl    0x804a080,(%esp)       # [MU]
 8048741: call   8048610                 # pthread_cond_wait(&CV, &
     MU)
          HANDLE_WAIT
 8048746: mov    0x804a060,%eax          # [COND]
 804874b: cmp     0x1,%eax               #
 804874e: setne  %al                     #
 8048751: test   %al,%al                 #
 8048753: STOP_RECORDING_READS
          GET_PARAMETERS
          jne    8048732                 # }
 8048755: STOP_IGNORING_WAITS
          ANNOTATE_WAIT
          movl    0x804a064,(%esp)       # [MU]
 804875c: STOP_RECORDING_READS
          STOP_RECORDING_WRITES
          call   80485c0                 # pthread_mutex_unlock(&MU)
```

Listing 6.1: Instrumented machine code with loops and condition variables

```
BEGIN_RECORDING_READS:
  delete old recorded read accesses of the current thread
  record all read accesses of the current thread up to now

STOP_RECORDING_READS:
  stop recording of read accesses

BEGIN_RECORDING_WRITES:
  delete old recorded write accesses of the current thread
  record all write accesses of the current thread up to now

STOP_RECORDING_WRITES:
  stop recording of write accesses

HANDLE_SIGNAL(cv):
  $TS_{cv}$ ← $max(TS_{cv}, TS_t)$ // saving the point of time
  for all recorded write accesses: // saving for data dependency analysis
    $TS_{cv}$(address) ← current thread segment $TS_t$
  $TS_t$ ← $NewSegment(TS_t)$ // signaling thread get new time

START_IGNORING_WAITS:
  inside_loop ← True

STOP_IGNORING_WAITS:
  inside_loop ← False

GET_PARAMETERS:
  Simulate loop body
  save parameters of simulated wait-call (cv,mu)

HANDLE_WAIT(cv,mu):
  if inside_loop:
    save parameters of wait-call (cv,mu)
  else: // recording synchronization
    dependency_found ← false
    thread segment $TS_w$ ← current thread segment $TS_t$
    for all recorded read accesses on address:
      if $\exists$ $V_{cv}$(Address): // data dependency detected
        $TS_w$ ← max($TS_w$, $TS_{cv}$(address))
        dependency_found ← true
    if dependency_found:
      current thread segment $TS_t$ ← thread segment $TS_w$
    else:
      current thread segment $TS_t$ ← max($TS_t$, $TS_{cv}$)

ANNOTATE_WAIT:
  thread segment $TS_w$ ← current thread segment $TS_t$
  for all recorded parameters (cv,mu):
    HANDLE_WAIT(cv,mu)
  delete recorded parameters
```

Listing 6.2: Overview of the functions used for Instrumentation process.

```
...
lock(mu);
BEGIN_RECORDING_READS();
BEGIN_RECORDING_WRITES();
...
STOP_RECORDING_WRITES();
signal(cv);
HANDLE_SIGNAL(cv);
...
STOP_RECORDING_READS();
STOP_RECORDING_WRITES();
unlock(mu);
...
```

Listing 6.3: Instrumented code of a signaling thread using condition variables

```
...
lock(mu);
BEGIN_RECORDING_READS();
BEGIN_RECORDING_WRITES();
...
GET_PARAMETERS();
while(condition) / for(;condition;) / if(condition)
{
    STOP_RECORDING_READS();
    START_IGNORING_WAITS();
    ...
    wait(cv, mu);
    HANDLE_WAIT(cv, mu);
    ...
}
STOP_RECORDING_READS();
STOP_IGNORING_WAITS();
ANNOTATE_WAIT(cv);
...
```

Listing 6.4: Instrumented code of a waiting thread using condition variables

The functions START_IGNORING_WAITS/STOP_IGNORING_WAITS inform the detector to ignore all **wait**-function calls within the **wait**-enclosing while loop. This is because within the while loop the synchronization is still not completely terminated. The detector ignores the **wait**-function calls only inside the loop, and will handle it after that by the injected functions.

Alternatively, we could ignore all **wait**-function calls in the program. But this might lead to missed synchronization operations that do not completely match the typical pattern of synchronization via condition variables. As an example, Figure 6.7 shows a program that uses *task queue* to distribute tasks between threads.

```
main() {
  create(&thread_1, &consumer);
  create(&thread_2, &consumer);
  ...
  create(&thread_n, &consumer);
  create(&thread_0, &producer);
}
```

```
void producer() {
  while(DATA) {
    Q.putTask();
    signal(cv);
  }
}
```

(a) Thread 1

```
void consumer() {
  while(TRUE) {
    wait(cv);
    do_something(Q.getTask());
  }
}
```

(b) Thread 2

Figure 6.7.: Using condition variables not following the standard pattern.

At the beginning, the consumer threads are blocked waiting for a signal *without* any condition. Then, the producer thread produces a task and wakes up an arbitrary thread that accomplishes the task. In this example, it is assured that the task is finished and a condition variable *cv* is only used to block the "idle" threads. If we ignore all **wait**-function calls, we would miss the resulted synchronizations and this case would not be treated properly.

6.6. Limitations

Lost signals and spurious wake ups are considered as main problems for hybrid race detectors. Static hybrid race detectors have to identify synchronization directly in program source code. They could use source code annotations based on the algorithm described in this chapter to mitigate these problems. For dynamic hybrid race detectors, our algorithm works fine and eliminates these problems.

For dynamic race detectors based on the lockset algorithm (Eraser-Style detectors), the partial order of events is not relevant at all; they report false warnings on lost signals anyway and miss races on spurious wake ups.

Pure happens-before race detectors additionally construct hb-relations based on lock/unlock operations. When using inter-thread event notifications via condition variables, accessing the variables of condition variables are within the lock-protected regions. Therefore, such detectors construct a hb-relation between lock-protected regions (see Figure 6.8).

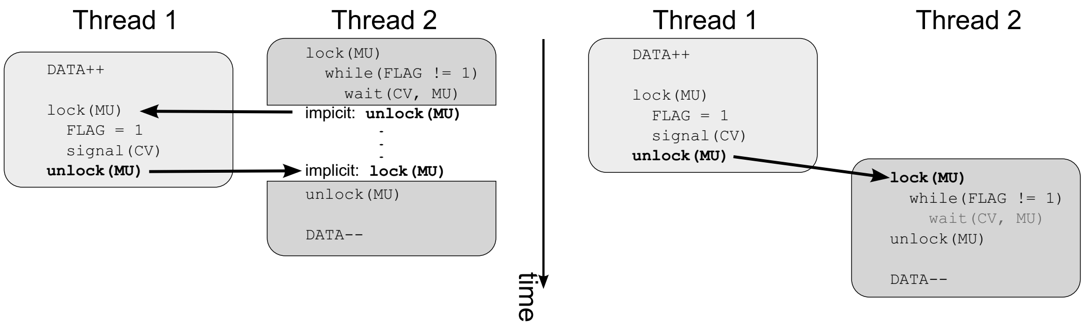


Figure 6.8.: hb-relation constructed from lock operations by a pure happens-before race detector.

As long as condition variables are used properly, hb-relations reflect the synchronizations correctly– the lost signal problem disappears. However, spurious wake ups still exist and the problem causes difficulties.

Nested Loop Conditions

Furthermore, there might be obscure use of inter-thread event notifications via condition variables defined by the programmer which do not match the common patterns. Our algorithm may not find the exact synchronization point in these cases.

As an Example, Listing 6.5 shows a synchronization with a condition variable using nested loops. The thread executes its own work in the outer loop. It is difficult to determine which ending point terminates the synchronization and in which loop level the synchronization occurs.

```
bool done = false;

while( !done ) {

  // Synchronization
  while( this ) {
    while( that ) {
      pthread_cond_wait(...);
    }
  }

  // working
  done = do_work();
}
```

Listing 6.5: Nested while-loops within a thread.

The algorithm presented has no special remedy for this case. The algorithm processes each loop as an ordinary loop. As long as the loop body does not seem too big for using a condition variable, each loop end is marked as a synchronization point. Only one of the marked synchronization points should be counted as a real ending synchronization point and the remaining are actually false synchronization points. The false synchronization points are considered several times and possibly too early by the detector. This causes a problem for the algorithm during the data dependency analysis only. If a synchronization point is considered too early, some data dependencies are not caught at runtime. Therefore the detector is not able to construct the actual hb-relation based on the required data dependencies. Ignoring the problem for data dependency analysis, hb-relations are constructed redundantly.

Nevertheless in the next chapter, we present a general approach to find different synchronization operations. It provide as solution that is able to identify programmer-defined synchronizations; among them such obscure synchronizations as nested loops.

Chapter 7.

Identifying Ad-hoc Synchronization

Programmers tend to implement their own synchronization primitives when available synchronization constructs are too slow. For instance, a programmer may write a spinning loop instead of using a library-supplied wait-operation. Furthermore, libraries may lack certain higher-level synchronization constructs such as barriers or task queues, forcing programmers to implement their own. We call such synchronization constructs implemented in programs *programmer-defined* or *ad-hoc*.

Ad-hoc synchronization operations occur surprisingly frequently. For instance, we found that eight of the 13 PARSEC benchmarks [3] contain between 32 and 329 ad-hoc synchronization code segments. For race detectors, ad-hoc synchronization presents a problem, in that race detectors are not aware of these constructs and thus generate an avalanche of false positives for them. For instance, Tian et al [44] observe an average of four million false positives generated for programs containing from 12 to 131 ad-hoc synchronization segments.

The subject of this chapter is the reliable detection and treatment of ad-hoc synchronization in race detectors, with the aim of eliminating false warnings.

We identified spin loops that check conditions as the basic pattern in ad-hoc synchronization. We provide a dynamic method that detects ad-hoc synchronization constructs reliably, provided they use spin loops that examine condition variables. The method dynamically and automatically identifies these loops by analyzing the object code. The signaling thread which signals the condition(s) (mostly by changing a flag) cannot be determined through analysis, but it can be found dynamically by instrumenting the code. Our method detects both reads and writes on condition variables and then establishes a hb-relation between signaling and signaled threads, thus preventing the generation of false warnings.

The method has been added to the race detector Helgrind$^+$[22, 21]. The results from substantial benchmark suites confirm that Helgrind$^+$ eliminates false warnings without missing true races.

7.1. Synchronization Operations

Dynamic data race detectors typically intercept calls to synchronization primitives in order to find ordering constraints. For instance, if a program calls the synchronization primitive *barrier_wait*, the detector intercepts this function call and records the fact that accesses after the barrier are not concurrent with accesses before the barrier. For a race detector to identify all races and to produce no false positives, it must be aware of the ordering effects of all synchronization constructs, including locks, monitors, signal/wait, conditions variables, spurious wakeup calls, barriers, etc. Our results given in Chapter 9 confirm that by tailoring the detection algorithm for each synchronization primitive, the detector extracts highly accurate ordering information, identifies all races, and keeps the number of false positives low.

But what if synchronization primitives are not supported by the detector? Then the usual detectors are not able to intercept them, know nothing about ordering effects, and therefore may produce numerous false positives. The number of false positives can be so high as to overwhelm the programmer.

The idea of this chapter is to identify a basic pattern that occurs in virtually all synchronization primitives and to extend the detection algorithm to handle this pattern. This pattern is the spinning read loop waiting for a condition to change. Once this pattern is handled well, we can, in fact, remove all code from the race detector dealing with synchronization primitives built upon this loop. Moreover, all synchronization libraries as well as ad-hoc synchronization based on the spinning read loop will be handled automatically, eliminating the need to enhance the detector for every library.

7.1.1. True and False Data Races

In addition to the data race classification given in Figure 3.5, we can classify false data races as follows:

- *apparent data races*, and
- *synchronization data races*

If a detector is not aware of a synchronization construct, it may report races where they are none, because they are actually prevented by the construct. Such cases are called *apparent* data races. Since detectors may not support all synchronization primitives, apparent races cause false positives. Figure 7.1 depicts a simple example that uses the synchronization primitive *barrier_wait(b)*. Assuming the primitive is unknown to the detector, it will report races regarding the variable DATA. The detector will consider all read operations after the barrier as races, although there is no concurrent write. If the synchronization primitive were known to the detector, the false races would disappear.

```
Thread 1, 2, ... n:
          lock(l)
            DATA++
          unlock(l)

            ...

          barrier_wait(b)

            ...

          print DATA
```

Figure 7.1.: Using the synchronization primitive *barrier_wait()* from an unsupported library causes apparent data races on DATA.

Figure 7.2 depicts a simple ad-hoc synchronization in which the second thread waits for condition variable FLAG. An uninformed standard detector would report an apparent race on variable DATA.

The second major reason for false positives are *synchronization data races*. Consider FLAG in Figure 7.2. Its accesses are unordered and constitute a real data race. However, this data race is a harmless and, in fact, *intentional*. The data race is necessary for proper synchronization. Intentional data races are often known as *synchronization races* [24, 44].

The example in Figure 7.1 also produces synchronization data races. To see why, consider the typical implementation of the barrier primitive in Figure 7.3. The intentional data races on the variable counter are harmless synchronization data races. Synchronization primitives require data races to enable competition for entering critical sections, for locking, for changing condition variables, etc.

Our aim is to refine Helgrind$^+$ in such a way that it does not report apparent or synchronization data races, while reporting all other races.

```
/* Initially FLAG is 0 */

...
```

```
DATA = 1

 ...

FLAG = 1
```

(a) Thread 1

```
while(FLAG != 1)
  /* do_nothing */

  ...

print DATA
```

(b) Thread 2

Figure 7.2.: Ad-hoc synchronization causes an apparent data race on DATA and a synchronization data race on FLAG.

```
lock(l)
  counter++
unlock(l)

  ...

while(counter != NUMBER_THREADS)
  /* do_nothing */

  ...
```

Figure 7.3.: Implementation of synchronization primitive *barrier_wait()* causes synchronization races on counter.

7.2. Ad-hoc Synchronizations

A good race detector should avoid false positives associated with ad-hoc synchronization and synchronization races. In this section, we propose a dynamic detection method based on the fact that the *spinning read loop* is the common pattern of almost all synchronization constructs, and, a major source of synchronization races. The method identifies spinning-loop synchronization correctly even if the spinning read is actually not entered (recall that programmers assume spinning loops are entered rarely). We discuss the underlying pattern first and present our detection algorithm afterwards. The algorithm identifies all spin-loop synchronization operations, including those in libraries.

7.2.1. Common Pattern in Ad-hoc Synchronization

The so called *spin-lock synchronization* is the most common and simplest synchronization construct [24]. It employs a *condition* shared by two threads. One thread executes a while-loop waiting for the value of the condition to be changed by the other thread. At the moment the value changes, the waiting thread is able to leave the loop and proceed. Figure 7.4 illustrates the use of the spin-lock. Thread 2 executes a spinning read loop on the shared variable `CONDITION`, until Thread 1 changes the value. The read and write operations are the source of a harmless synchronization data race that need not be reported to the user.

```
    ...

do_before(X)

set CONDITION to TRUE
       .
       .
       .
```

(a) Thread 1

```
while(!CONDITION){
  /* do_nothing */
  }

do_after(X)
     .
     .
     .
```

(b) Thread 2

Figure 7.4.: Spinning read loop pattern.

A number of publications analyzed different implementations of synchronization operations [28, 29, 17, 24, 44] and observed that the *spinning read loop* is a common pattern used for implementing synchronization constructs. For example, the implementation of the barrier primitives in Figure 7.3 also uses a spinning read loop on a flag with a *counterpart write* ending the spin.

Another example is different lock implementations in various thread libraries that use the *Test-and-Set* or compare and swap construction. Test-and-Set is also used by the PThread library to implement the lock primitives as shown in Figure 7.5.

By executing the **lock**`(l)` operation, each thread executes an atomic Test-and-Set instruction i.e. `Test-and-Set(lock_flag)`. This instruction reads and stores the value of `lock_flag` and sets the variable `lock_flag` to `TRUE`. If the `lock_flag` is available and not acquired by other threads, then the Test-and-Set instruction returns `FALSE`, that lets the executing thread wins to enter the critical region. Other threads have to wait and spin on the `lock_flag` (**`while`** `(lock_flag);`) until there is a possibility that Test-and-Set instruction can succeed.

```
 1 bool lock_flag = FALSE;
 2
 3 lock() {
 4   while (Test-and-Set(lock_flag) {
 5     while(lock_flag);
 6   }
 7 }
 8
 9 unlock() {
10   lock_flag = FALSE;
11 }
```

Figure 7.5.: Implementation of lock/unlock operations in various libraries (e.g. Pthread library).

To avoid executing the Test-and-Set instruction repeatedly, threads do spinning read on line 5. Running the Test-and-Set instruction repeatedly causes cache invalidations and generating cache coherence messages. This could lead to significant overhead. In the implementation, there is a spinning read loop on variable `lock_flag` at line 5 (loop: `while(lock_flag);`). Along with the counterpart write at line 10 (`lock_flag = FALSE;`), a synchronization data race on variable `lock_flag` is created.

One can see that the atomic instruction `Test-and-Set(lock_flag);` in line 4 simultaneously reads and writes the `lock_flag` variable, and consequently causes synchronization data races with lines 5, 10, and itself.

7.2.2. Detecting Spinning Read Loops

In the previous section, we found that the spinning read loop and its counterpart write are the common construct for ad-hoc synchronizations. Helgrind$^+$ finds spinning read loops in program code and instruments them before runtime. When the program is executed, Helgrind$^+$ establishes the correct happens-before relations between spinning read loops in one thread and counterpart writes in another thread, so that the detector is aware of synchronizations.

The general idea of our method is as follows: Helgrind$^+$ searches the binary code to find all loops. This is done by building a blockwise control flow graph on-the-fly before the actual runtime. Next, it narrows the set to spinning read loops based on the following criteria:

- Evaluating the loop condition involves at least one load instruction from memory.

- The value of the loop condition is not changed inside the loop body.

We instrument each spinning read loop and mark the variables that affect the value of the loop condition (this may be a single flag or several variables, if the condition is an expression). However, at this point we do not know where the counterpart write is. Next, we discuss the instrumentation that finds the write.

Being a runtime race detector, Helgrind$^+$ monitors all read/write accesses. The state of a variable indicates whether it is used by one thread only (state *exclusive*) or multiple threads (state *shared*). When entering a spinning loop, the states of the variables affecting the loop condition are set to a special state called *spin*. Two cases are possible:

(a) In the first case, the counterpart write does not happen before entering the spin loop. In this case, Helgrind$^+$ waits for the first write operation that affects the loop condition. If the write operation is performed by another thread other than the spinning thread, then this is the counterpart write. When leaving the loop, Helgrind$^+$ records a happens-before edge between the spinning read loop and its counterpart write.

Consider the example in Figure 7.4. `CONDITION` is the condition variable in the spinning read loop. Assume the spinning read is entered by Thread 2 before the counterpart write. The happens-before relation is constructed based on the data dependency (*write/read* dependency) on `CONDITION`. Thus, the detector will be aware of this synchronization operation and no race will be reported on `X`. The warning regarding the benign synchronization race on `CONDITION` is also suppressed, since it is marked as being in the special state *spin*.

(b) In the second case the counterpart write happens first: One or more variables affecting the loop condition are written before the loop is entered. Helgrind$^+$ sets the states of the changed variables to *exclusive* and records the location of the write instructions. As soon as the instrumented spinning read loop is entered, the detector notices that variables affecting the loop condition have been changed. The loop itself terminates immediately. Helgrind$^+$ records the hb-relation as before and sets the states of the changed variables to *spin*. In this case no actual spinning happens—the loop condition is evaluated only once and the loop ends.

The method described by Tian et al [44] fails to establish the hb-relation in the second case, because this method relies on the loop spinning several times. It also fails to recognize inter-thread event notifications (*signal()/wait()*) in case of lost signals. A signal is lost if a thread sends a signal before any other thread waits for it. In this case, no spinning read happens at runtime. The *wait()* primitive is not even executed, since the condition variable is already set earlier by the signaling thread; the loop terminates immediately. Failing to take signaling into account may lead to false positives. Another tricky case involves spurious wake ups. These can lead to false negatives (missed races) when threads uses same condition variables (as discussed in Chapter6). Helgrind$^+$ handles all of these cases correctly.

Another problem by the method of Tian et al [44] concerns the heuristic of using a threshold iteration count in order to distinguish spinning read loops from ordinary loops. If the spinning read loop dose not spin long enough to reach the threshold value, the detector misses the spinning read loop and generates false positives. On the other hand, if the threshold value is too low, ordinary loops in the program could be mistaken for spinning read loops, which also results in missed races. Thus, without exploiting the semantics information by dynamic code analysis just before runtime, one may easily miss synchronizations or misinterpret them, since actual spinning reads may not happen at all at runtime or might not reach a preset threshold value.

7.2.3. The Algorithm

Conceptually, our method is divided into three phases:

- pre-instrumentation phase,
- instrumentation phase, and
- runtime phase

In the pre-instrumentation phase, loops in the program are recognized and then only the spinning read loops are selected to be instrumented just-in-time in instrumentation phase. During runtime, a runtime data dependency analysis is carried out to construct the hb-relation between the related parts (spinning read loop and counterpart write).

Recognizing loops in the program is performed by means of online control flow analysis. We construct a blockwise control flow graph on-the-fly based on the current superblock (a superblock contains a maximum of three basic blocks) and consider loops with three to seven basic blocks in the graph. We check whether they are spinning read loops or not. In our experiments, we found three to seven basic blocks deliver good results, since the spinning read loops are typically small loops with few instructions. Decreasing this number may result in missing some spinning read loops and producing some false positives.

On the other hand, increasing the number of basic blocks causes additional overhead.

Figure 3 provides a high level description of spinning read loop detection algorithm. The first step constructs the data dependency table D_l for every loop l. $D_l(condition_l)$ returns all variables that the loop condition $condition_l$ depends on within the loop l. The data dependency analysis takes function calls into account. Step 2 determines for all variables v that the $condition_l$ depends on, whether v is modified inside the loop. If there is an assignment to any such v, then the loop l is not a spinning read loop. Otherwise, the loop is marked as performing spinning reads only, and the variables of $D_l(condition_l)$ are prepared for instrumentation.

```
foreach loop l do

    1. $D_l(condition_l)$: the set of variables, on which the condition
       $condition_l$ of the loop $l$ depends

    2. foreach $v \in D_l(condition_l)$ do
         if $(D_l(v) \neq \emptyset)$ then        // $v$ is modified
           return;
         end
       end

    3. mark $l$ as spinning read loop
       prepare all $v \in D_l(condition_l)$ for instrumentation

end
```

Algorithm 3: Basic algorithm for detecting spinning read loops.

Loops are converted to conditional branches at low level code. Hence, we must consider all conditional branches in the low level code. We search the control flow graph for loops that span a maximum number of three to seven basic blocks. Then, we track the dependencies of each variable within these basic blocks by constructing a data dependency table. The data dependency table is built up with respect to registers at the low level code. All temporary variables and addresses in basic blocks are traced to identify the registers they depend on. By means of the dependency table we can now check if the loop condition variable depends on a register that is the target of a *load instruction.* If the load addresses stay constant, then the loop condition variable is not modified within the loop body and the loop is marked as a spinning read loop. We instrument the spinning read loop and insert the required instructions to intercept and analyze it at runtime.

```
volatile int spinlock = 1;

//worker thread
void *run_pm(void* arg) {
  spinlock = 0;
  return NULL;
}

//parent thread
int main(){
  pthread_t threadid;
  pthread_create(&threadid, NULL, &run_pm, NULL);

  while (spinlock != 0) {
    sched_yield();
  }

  pthread_join(threadid, NULL);
  return 0;
}
```

Figure 7.6.: An example for spinning read loop.

Figure 7.6 shows exactly how the algorithm works in more detail: The spinning read loop is located in a parent thread. In low level code, the spinning loop is converted to conditional jumps resulting in the object code depicted in Figure 7.7. For the sake of simplicity, we just show the code for the spinning read loop part: `while (spinlock != 0)`. The first instruction is a `mov` operation that loads a value from memory to register `eax`. Then the conditional jump is evaluated. Register `eax` stays constant within the loop body, since it depends on itself only. Thus, the loop is marked and instrumented as a spinning read loop.

As in the previous chapter, we use the functions BEGIN_RECORDING_READS and STOP_RECORDING_READS to record all read accesses at this specific code region. They will be used during runtime analysis for extracting write/read dependencies and constructing hb-relations.

Unlike the previous example, the example provided in Figures 7.8 and 7.9 is a non-spinning read loop. It is a small while-loop, but the variable `runp` is modified within the loop body. This can be easily derived from analyzing the object code given in Figure 7.9. Register `rdx` is loaded from a memory location

```
        BEGIN_RECORDING_READS
4006a0: mov    0x200462(%rip),(%eax) # eax <- load(...); [spinlock]
4006a6: test   %eax,%eax             #
4006a8: jne    40069b                # while(spinlock) {
        STOP_RECORDING_READS
40069b: call   400548                # sched_yield()
40069e: cmpb   0x0,0x200462          #
4006ab: je     40069b                # }
```

Figure 7.7.: Machine code of spinning read loop depicted in Figure 7.6.

by instruction `mov (%rax),%rdx` and then the loop condition is evaluated at the Line `e44ed` by `test %rdx,%rdx`. However, `rdx` is modified within the loop body by the instruction `mov %rdx,%rax` at Line `e44e7`. In fact, the loading address (`rax`) of the load instruction `mov (%rax),%rdx` is modified.

```
while (runp != NULL)
  if (runp->dso_handle == dso_handle)
    break;
  else
    {
      lastp = runp;
      runp = runp->next; // loop variable ''runp'' is modified!
    }
```

Figure 7.8.: An example for non-spinning read loop.

```

e44e1: cmp    %r9,0x20(%rdx) #
e44e5: je     e44f4          # if (runp->dso_handle == dso_handle)

e44e7: mov    %rdx,%rax      # rax <- rdx; modified!

e44ea: mov    (%rax),%rdx    # rdx <- load(rax); [runp]
e44ed: test   %rdx,%rdx
e44f0: jne    e44e1          # while (runp != NULL)
```

Figure 7.9.: Machine code of non-spinning read loop depicted in Figure 7.8.

7.2.4. Limitations

There could be obscure implementations of spinning read loops that do not follow the common patterns. As an example, the code presented in Figure 7.10

is a type of spinning loop synchronization done by a **`do-while`** loop. However, the loop condition variable `c` inside the loop body is changed in each iteration and based on our algorithm the detected loop is not instrumented as a spinning read loop.

```
static int COND = 0;

void Thread1() {
  X = 1;

  lock(l);
    COND = 1
  unlock(l);
}

void Thread2() {
  int c;
  do {
    lock(l);
      c = COND; // loop condition changed!
    unlock(l);
    sleep(1);
  } while(c);

  if (c) {
    X = 2;
  }
}
```

Figure 7.10.: Obscure implementation of a spinning read loop.

7.3. Universal Race Detector

A side benefit of the method mentioned in this chapter is that it can also be applied to unknown libraries. Helgrind$^+$ currently uses information about the synchronization constructs of PThreads, but if application programmers use different libraries, then our enhanced Helgrind$^+$ can also detect races reliably, provided the libraries are based on spin loops. Note that even operating system calls such as *wait* that relinquish the processor are typically used inside loops and therefore detectable by Helgrind$^+$. A surprising result is that information about PThreads can be removed entirely from Helgrind$^+$, resulting in only a minor increase in false positives. Thus, Helgrind$^+$ with spin loop detection can be seen as a *universal race detector* that is aware of synchronization operations.

We overcome the serious limitation of prior work which limits detectors to only synchronization primitives of a particular library.

7.3.1. Detecting Synchronization Primitives

The above method is a general approach capable of detecting synchronization operations implemented by spinning read loops e.g. locks, barriers, etc.

A race detector based on this approach is a pure happens-before detector. It cannot make use of the lockset algorithm, because it is not aware of locks. In our case, if we turn off the support of Pthreads so that synchronization primitives of Pthreads are not directly intercepted, we get a pure happens-before detector. Our empirical results show that relying only on this general approach for identifying synchronization operations might be too limited in some situations. There may be obscure implementations of spinning read loops that are difficult to detect, leading to false positives. We will show that using this approach as a complementary method achieve better results.

Our method is based on dynamic binary instrumentation. It does not need any program source code or user intervention such as source code annotations and therefore is non-intrusive in the build process. The semantic analysis of the program code is done automatically during just-in-time binary instrumentation.

Chapter 8.

Implementation of Helgrind$^+$

This chapter describes the main points in the implementation of Helgrind$^+$. We provide a general overview and discuss important data structures and algorithms used in the Helgrind$^+$ implementation.

8.1. Valgrind Framework

Valgrind [33, 31, 32] is a framework for dynamic binary instrumentation. It consists of a virtual machine (core) and different tools (skins) built on top. The tools perform various analyses of program execution using instrumentation. Some existing tools built on top of Valgrind are, for example, *Memcheck*[1] or *Cachegrind*[2], etc. Our tool Helgrind$^+$ is built on top of Valgrind, too.

The virtual machine disassembles binary code into a platform independent code called *Intermediate Representation* (IR): IR is a RISC-like intermediate language that uses main memory and registers with some temporary variables as memory locations. A tool inserts its instrumentation functions into the IR code. The instrumented IR code is handed back to the virtual machine. The virtual machine *resynthesizes* the machine code from the instrumented IR. Then, the code is executed. During execution, the instrumentation functions are also executed and tools can monitor the memory and register contents for their analysis. This process is depicted in Figure 8.1.

Valgrind performs instrumentation just-in-time and on demand. It is blockwise based on Valgrind *superblocks*. A superblock contains a maximum of three basic blocks. A *basic block* is a sequence of IR instructions ending with a branch operation (control flow branch). Therefore, a superblock could have more than one exit point (end of each basic block could be an exit point for the superblock).

[1]A memory checker tool.
[2]A cache profiler tool.

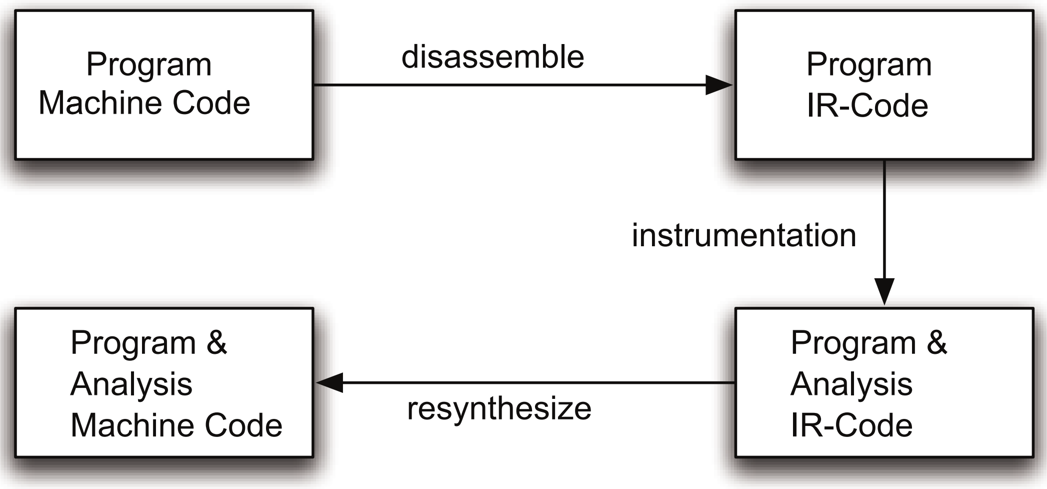


Figure 8.1.: Instrumentation process in Valgrind.

If a superblock is going to be executed, then superblock instrumentation is enforced. Superblocks not executed , are not instrumented. The superblock is instrumented only once and instrumented superblocks are stored for possible reuse during program execution. However, it is possible that an instruction appears in different superblocks and thus the instruction could be instrumented several times. The instrumentation has no side effects on program functionality. Each instruction is always instrumented the same way regardless of its context.

A superblock could have entry points within its body. In this case, only the part following the entry point will be instrumented. Superblocks are allowed to overlap with each other.

Valgrind supports multi-threaded programs and schedules threads in a fine-grained way. It is able to intercept function calls from the Pthread library. The virtual machine is single-threaded. Threads are created and Valgrind cares for very fine granular interleavings, but at any point of time, only one thread will be executed.

Event Handler and User Requests In addition to instrumentation of IR, Valgrind proposes two other mechanisms to get more information about the program execution: Many events, such as memory allocation, or lock acquisition are intercepted by Valgrind and could be delivered to the tool by an *event handler*. Also, a tool can define *user requests*; functions in the program that are going to be analyzed by the tool, so that Valgrind diverts the program control to a function provided by the tool – annotations in program source code are implemented in this way.

All together this functionalities makes Valgrind a powerful and flexible tool for all kinds of runtime checking.

8.2. Machine Code and Intermediate Representation (IR)

It is difficult to describe or locate high level constructs such as loops directly in IR code, because IR is relatively expanded compared to machine (object) code. It could be useful for finding such constructs, if we consider *machine code* first, instead of IR code. But instrumentation in IR code is easier. The operations in IR are more explicit: Each machine instruction is expanded into several simple IR instructions. There are less varied IR instructions and therefore the instrumentation process in IR level is easier.

Figure 8.2 shows different code representations of a sample program. The example depicts a small program in *C/C++* source code and its *x86* machine code. The machine code uses instructions for arithmetic operations, branch operations and function calls. At lines 8 and 9 of the machine code, the **`if`** construct is replaced by a "conditional jump" and "compare" instructions. At lines 11 to 14, first the parameters of a function are pushed onto the stack, before jumping to the function address located at `0x8048330`. How the parameters are delivered after that, depends on the calling conventions which differ for each platform or hardware architecture. It is very difficult to infer that the address at `0x8048330` is a function call to `printf`. In machine code, there are no explicit write/read instructions. Lines 4 and 8 demonstrate different instructions to access variables.

The Intermediate Representation abstracts from different machine instructions and forms a platform independent representation of program code. Figure 8.3 shows the intermediate representation of the example given in Figure 8.2. The IR code of this program is divided into two superblocks including their Instruction Marks (IMarks). Each IMark relates the IR instructions to its specific machine code instruction, as every instruction in machine code is translated into several IR instructions.

At the end of each superblock, there is a pointer to the next superblock according to the execution order. A *conditional* jump cannot be used as the last instruction in a superblock. Instead, a *goto* statement is used at the end of each superblock,

```
int a = 2;
int b = 2 * a;

if(!b) return;
printf("%d", b);
```

(a) Source code

```
int a = 2;
 8048405:  movl   0x2,-0x8(%ebp)
int b = 2 * a;
 804840c:  mov    -0x8(%ebp),%eax
 804840f:  add    %eax,%eax
 8048411:  mov    %eax,-0xc(%ebp)
if(!b) return;
 8048414:  cmpl   0x0,-0xc(%ebp)
 8048418:  je     804843a
printf("%d", a);
 804841a:  mov    -0xc(%ebp),%eax
 804841d:  mov    %eax,0x4(%esp)
 8048421:  movl   0x8048510,(%esp)
 8048428:  call   8048330
```

(b) Machine code

Figure 8.2.: Example for machine code generated for x86.

as shown in the Figure 8.3. Thus, within a superblock there are no *unconditional* jumps (e.g. *goto* statement). If an unconditional jump is encountered, the superblock ends, and a pointer to the new superblock is set. As an alternative, *jump chasing* could be used so that the jump address is evaluated, and the code at the jump destination will be attached to the current superblock. Within a superblock, temporary variables `tn` are created that follow the *single static assignment* rule (only one assignment to each temporary). Temporary variables, called *temporaries*, are used only as arguments for IR operations, and work according to the load/store principle.

```
 1 ---- IMark(0x8048405, 7) ----
 2 PUT(60) = 0x8048405:I32
 3 t35 = Add32(t30,0xFFFFFFF8:I32)
 4 STle(t35) = 0x2:I32
 5 ---- IMark(0x804840C, 3) ----
 6 PUT(60) = 0x804840C:I32
 7 t37 = Add32(t30,0xFFFFFFF8:I32)
 8 t39 = LDle:I32(t37)
 9 ---- IMark(0x804840F, 2) ----
10 t16 = Shl32(t39,0x1:I8)
11 PUT(0) = t16
12 ---- IMark(0x8048411, 3) ----
13 PUT(60) = 0x8048411:I32
14 t40 = Add32(t30,0xFFFFFFF4:I32)
15 STle(t40) = t16
16 ---- IMark(0x8048414, 4) ----
17 PUT(60) = 0x8048414:I32
18 t43 = Add32(t30,0xFFFFFFF4:I32)
19 t22 = LDle:I32(t43)
20 IR-NoOp
21 PUT(32) = 0x6:I32
22 PUT(36) = t22
23 PUT(40) = 0x0:I32
24 PUT(44) = 0x0:I32
25 ---- IMark(0x8048418, 2) ----
26 PUT(60) = 0x8048418:I32
27 t52 = CmpEQ32(t22,0x0:I32)
28 t51 = 1Uto32(t52)
29 t50 = t51
30 t53 = 32to1(t50)
31 t45 = t53
32 if (t45) goto {Boring} 0
       x804843A:I32
33 goto {Boring} 0x804841A:I32
```

(a) IR Superblock 1

```
 1 ---- IMark(0x804841A, 3) ----
 2 t7 = GET:I32(20)
 3 t6 = Add32(t7,0xFFFFFFF4:I32)
 4 t8 = LDle:I32(t6)
 5 PUT(0) = t8
 6 ---- IMark(0x804841D, 4) ----
 7 PUT(60) = 0x804841D:I32
 8 t10 = GET:I32(16)
 9 t9 = Add32(t10,0x4:I32)
10 STle(t9) = t8
11 ---- IMark(0x8048421, 7) ----
12 PUT(60) = 0x8048421:I32
13 STle(t10) = 0x8048510:I32
14 ---- IMark(0x8048428, 5) ----
15 PUT(60) = 0x8048428:I32
16 t12 = Sub32(t10,0x4:I32)
17 PUT(16) = t12
18 STle(t12) = 0x804842D:I32
19 ---- IMark(0x8048330, 6) ----
20 PUT(60) = 0x8048330:I32
21 t4 = LDle:I32(0x804A008:I32)
22 goto {Boring} t4
```

(b) IR Superblock 2

Figure 8.3.: Example of intermediate representation (IR-code).

There are specific IR operations that are classified into the following categories:

Arithmetic operations are used for calculation and comparison as well as for address calculations. In the first superblock in Figure 8.3, there are some examples at lines 3 and 27. The suffix `32` indicates that the operands are 32-bit long. Operands are temporaries only and the result is also assigned to temporaries (max. four operands).

Accesses to registers of the processor are done by `PUT` and `GET` commands.

The actual registers are represented by an architecture specific constant offset number. For example in x86 the program counter has the offset `60`, the accumulator has the offset `eax` and `0`, and the stack pointer has the offset `16`. Examples for them are on lines 2 or 21 to 24 in the first superblock.

Access to main memory is performed explicitly by `ST` or `LD` instructions. When using these instructions, the most significant bit (MSB) and the loading or storing word length is needed, too. Temporaries are allowed as address (operand) only. See line 4 or 8 in IR Superblock 1.

Conditional branches are used in the form of **`if`**`(tn){info}` **`goto`** `y`, where `tn` is a 1-bit temporary variable and `y` is the branch destination address. `info` gives some additional information about the branch operation. For example, on line 32, one can find such a branch operation. In this case, the keyword `Boring` indicates the branch as an ordinary branch operation.

Dirty Helper provides function calls for the tool. It is not trivial to implement complex functions in IR code. The dirty helper commands call *C* functions inside the tool so that the programmer comfortably can use *C* language syntax for analysis functions (see the example given in Figure 8.4).

Additional information is the IR instructions that do not affect the program behavior. However, they provide useful information about the code for the instrumenting tool, such as, *IMarks* that provide the address and the length of machine instruction.

A Valgrind tool is allowed to arbitrarily modify the code. Of course, the program semantics should not be affected. For instance, if the Helgrind$^+$ tool wants to monitor all read/write accesses, it has to insert its own analysis code before or after each load or store operation. Figure 8.4 provides an example: A simple tool counts the number of jump conditions evaluated to `true`. What we need is to insert an analysis function by the *Dirty Helper* command before each jump instruction and pass the condition temporary as a parameter to the function. The function represented by *Dirty Helper* (line 32) checks the value of the condition at runtime and increments a counter, if appropriate. Line 32 depicts the function inserted by *Dirty Helper*.

As with machine code, IR function calls are not easy to identify. Additionally, loading dynamic functions is platform dependent and varies for each platform. Valgrind's workaround for this problem is to overwrite some specific library functions with help from the operating system[3]. Thus, by calling a library

[3]LD_PRELOAD environment variables.

```
PUT(60) = 0x8048418:I32
t52 = CmpEQ32(t22,0x0:I32)
t51 = 1Uto32(t52)
t50 = t51
t53 = 32to1(t50)
t45 = t53
DIRTY 1:I1 ::: countTrueConds{0x38019600}(t45)
if (t45) goto {Boring} 0x804843A:I32
goto {Boring} 0x804841A:I32
```

Figure 8.4.: Instrumented IR code.

function, a Valgrind function is called instead of the actual library function. Valgrind receives the function calls together with parameters at runtime.

8.3. Shadow Memory

Shadow memory [33] is used by many tools including Helgrind$^+$. Using shadow memory, a tool manages the meta data (e.g. vector clock of last access) about each word of memory which will be used for program analysis. The program itself, which is going to be analyzed by the tool, does not know about the shadow memory.

The implementation of shadow memory in Helgrind$^+$ is done by a two stage table, similar to page tables in virtual memory management. Assume *addr* is a memory address in the user program. To determine the related shadow value of *addr*, a part of *addr*, exactly the high level 16 bits are considered as an index for the first table, that points to a second table. In the second table the low 16 bits of *addr* are used as an index to locate the actual shadow value. The second table will be allocated if it is needed. This simple structure facilitates some optimizations, i.e. fast access to multi byte operations and compact shadow values.

Furthermore, there are shadow registers for hardware registers available to Helgrind$^+$ similar to other tools. They are used for internal purposes and can be read and written only by tools. Valgrind keeps the registers and shadow registers during thread switches. Tools are able to access only the shadow values of the current active thread.

Helgrind$^+$ has to maintain state information not only for each thread, but also for each data used by the program. As Helgrind$^+$ is unaware of high level data structures, it can only operate on memory locations with byte level granularity.

State information about memory locations is stored with shadow memory. For every byte of memory used in the program, a shadow word is stored. Valgrind provides an extra memory pool for the shadow memory and other data structures such as mutex or thread segment information so that Helgrind$^+$'s data will not mix with the data of the program. Tools may use 32-bit shadow words, which is too small for our race checker, so we use 64-bit shadow memory. We explain the structure of 64-bit shadow word used by Helgrind$^+$ in the following sections.

8.3.1. Shadow memory for MSM-long

Our extended memory state machine for long-running applications, *MSM-long*, has eight different states. In the state *Shared-Modified2*, both lockset and thread segment information is stored. Therefore, Helgrind$^+$ uses a 64-bit shadow word for each memory location. Figure 8.5 demonstrates the structure of the 64-bit shadow words in different states.

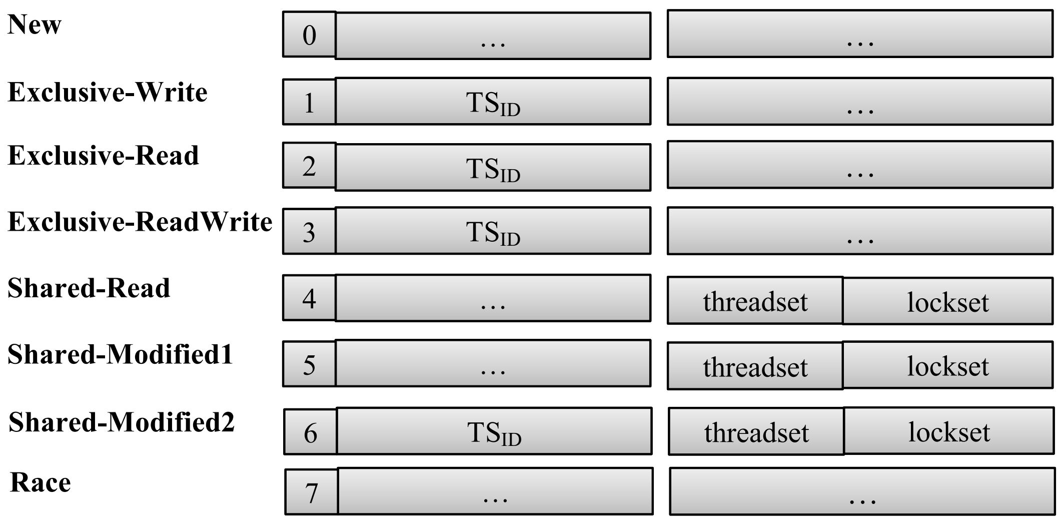


Figure 8.5.: Structure of 64-bit word shadow value and state encoding for MSM-long

The three most significant bits of the 64-bit shadow value are used to encode the state (0 - 7). The interpretation of the other bits depends on the state. Three bytes are used to store the thread segment identifiers in exclusive states and *Shared-Modified2*. In the exclusive states, the second 32-bit word is unused. Lockset information is stored in the four least significant bytes (second 32-bit word) of shared states. The states *New* and *Race* use three bits only to store encoded state. If a memory location is in an exclusive state, the thread segment ID of the last access is stored. The lockset is not initialized unless one of the shared states is reached.

Moreover, a fixed heap provides structures that require dynamic allocation during runtime to avoid dynamic memory allocation. These structures include condition variable and mutex information, as well as thread segments. Reference counting is used to determine which segments are free and is no more referenced; a simple garbage collection algorithm returns unused segments back to the heap.

8.3.2. Shadow memory for MSM-short

As Figure 8.6 shows, Helgrind$^+$ takes advantage of 64-bit shadow values for memory states for short-running applications. Bits are used differently in different states. The state is stored in the first three bits. Not all information is relevant in each state. Threadsets, for example, are not tracked in exclusive states, so we do not need to store them in these states. A thread segment identifier is stored in the remaining bits of the first 32-bit word; candidate lockset and threadset are stored in the second 32-bit word.

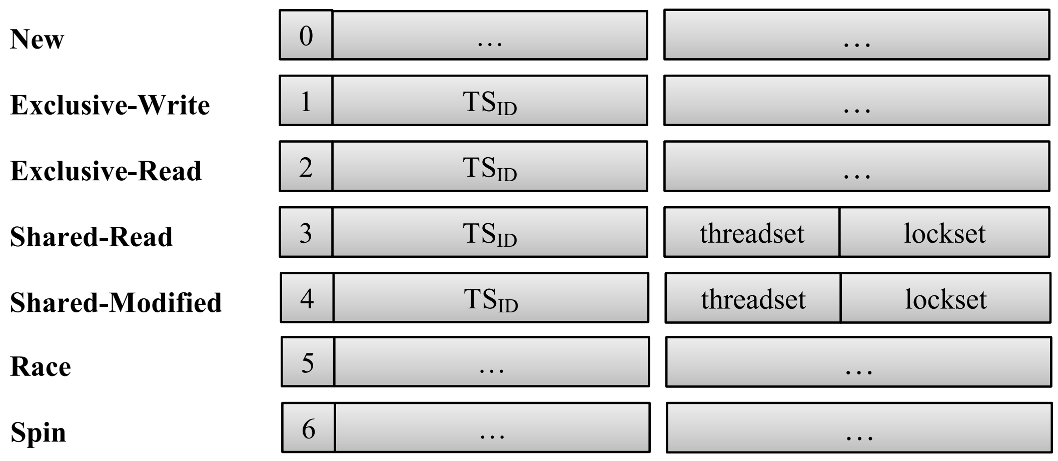


Figure 8.6.: Structure of 64-bit word shadow value and state encoding for MSM-short

We use the state *Spin* for the spinning read loop detection and assign it to the special value of *spin*, if the variable is used for spinning read loops (similarly, the state *Spin* is used for shadow memory of MSM-long).

Shadow values in Helgrind$^+$ double the space overhead compared to tools that use 32-bit shadow value. But there are several optimizations that could reduce the memory overhead in Helgrind$^+$, as 64 bits are needed in only two states.

8.4. Control Flow Analysis

A challenge during the instrumentation process is the control flow dependency that results in *Control Flow Graphs* (CFG) for identifying loops. The control flow analysis is done on the fly and based on Valgrind's suprblocks in IR. A superblock includes up to three basic blocks and contains a `next` pointer to the successive superblock. Superblocks symbolize the nodes with a maximum outdegree of four. Normally, loops are identified by finding a cycle within the CFG. This is done by a depth-first search with the limited depth of `n` (number of superblocks), which could be set by the command line *–CFG = n* in Helgrind$^+$.

Valgrind instruments a program block by block at runtime. We construct the control flow graph on the fly based on the executing superblocks and analyze their instructions. Thus, the complete control flow graph for the whole program is not available. Since loops are converted to *conditional branches* in IR code, we have to analyze all branches at IR level to find loops.

8.4.1. Branch Analysis

As each branch could belong to a loop, we have to check all of them in IR. At first, we determine the jump destination addresses to identify the loop body. Figure 8.7 shows a branch that is part of a while-loop in machine code and the corresponding IR. We have to search for a conditional branch at the end of the superblock, because Valgrind ends superblocks directly after an IR branch.

In Figure 8.7, the conditional branch in machine code at line 6 corresponds to the conditional branch in IR-Code at line 19. Surprisingly, the destination address is not `0x8048732` as in machine code, but `0x8048755`. This is because Valgrind could invert the conditional branches. For this reason, the condition of IR branch is inverted and the destination jump address is exchanged with the address of next superblock (`0x8048755`). In the example provided, you find the absolute jump destination address `0x8048732` at line 20.

Generally, there are two kinds of IR branches: *direct* an *inverted.* Figure 8.8 compares them. The difference is that in a inverted branch the destination address of the jump instruction (**`if`** `(t5)`**`goto`**) is the `next` pointer (beginning

```
8048732: # First instruction
    of loop body
...
804874b: cmp    0x1,%eax
804874e: setne  %al
8048751: test   %al,%al
8048753: jne    8048732  <--
8048755: # First instruction
    after the loop
```

(a) Machine cod of a conditional branch.

```
...
[ before is register eax used
    an in t8 stored.]
---- IMark(0x804874B, 3) ----
IR-NoOp
---- IMark(0x804874E, 3) ----
t1 = CmpNE32(t8,0x1:I32)
t2 = 1Uto8(t1)
PUT(0) = t2
---- IMark(0x8048751, 2) ----
PUT(32) = 0xD:I32
t3 = 8Uto32(t2)
PUT(36) = t3
PUT(40) = 0x0:I32
PUT(44) = 0x0:I32
---- IMark(0x8048753, 2) ----
PUT(60) = 0x8048753:I32
t4 = And32(t3,0xFF:I32)
t5 = CmpEQ32(t4,0x0:I32)
if (t5) goto 0x8048755:I32
goto 0x8048732:I32
```

(b) IR code of the same conditional branch (the code is simplified for more readability).

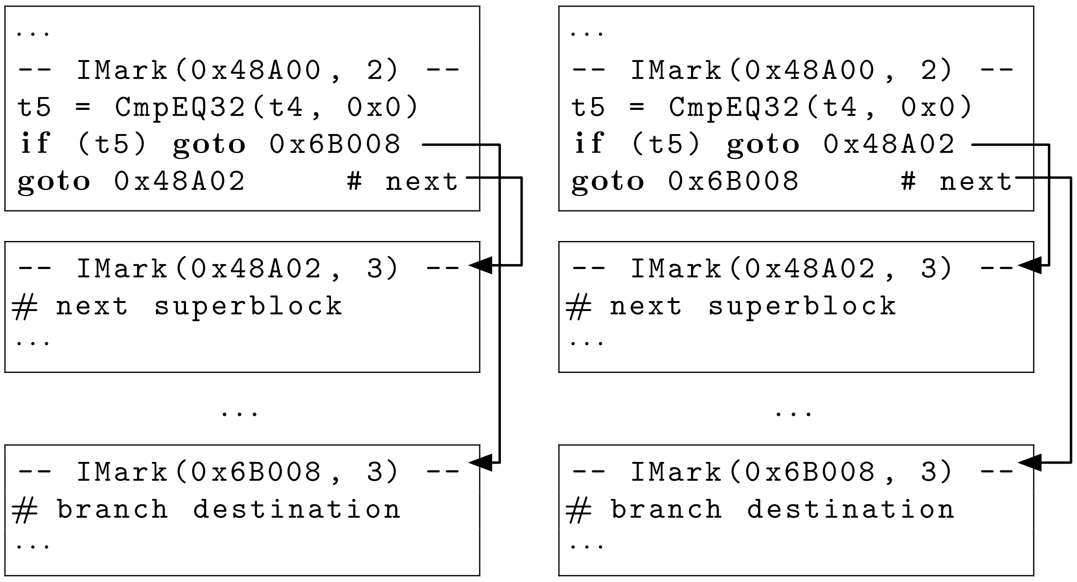
Figure 8.7.: A conditional branch in IR.

address of the next superblock). The reason of having inverted branches is the limitation of IR instructions that not every condition could be directly formulated. There is no specific command for testing an *unequal condition* (e.g. `jne` as in machine code). Valgrind negates the condition and exchanges the jump destinations.

Overall, the absolute destination address and the branch distances are determined in both cases . If the branch is not taken the execution continues to the next instruction right after the branch instruction. It is also possible to calculate the address of this instruction. The loop body is located between these two addresses.

8.5. Instrumenting Loops in IR

For each conditional branch at the IR level, we have to examine if they are part of synchronization with inter-thread event notification (condition variables) or belongs to a spinning read loop.



```
...
-- IMark(0x48A00, 2) --
t5 = CmpEQ32(t4, 0x0)
if (t5) goto 0x6B008
goto 0x48A02       # next

-- IMark(0x48A02, 3) --
# next superblock
...

...

-- IMark(0x6B008, 3) --
# branch destination
...
```

```
...
-- IMark(0x48A00, 2) --
t5 = CmpEQ32(t4, 0x0)
if (t5) goto 0x48A02
goto 0x6B008       # next

-- IMark(0x48A02, 3) --
# next superblock
...

...

-- IMark(0x6B008, 3) --
# branch destination
...
```

Figure 8.8.: IR branches: left side *direct* branch and and right side *inverted* branch.

We applied our algorithm explained in Section7.2.3 for spinning read loop detection to all temporaries, addresses, and registers in IR. We considered loops up to seven Valgrind's superblocks. If the branch condition (a condition temporary e.g. `t5`) depends on a register that is target of a load instruction (this could be done by the `Ld` instruction) and the loading address stays constant within the loop body, this branch is a candidate for a spinning read loop. Then, it is instrumented and the instrumenting function is inserted by a Dirty Helper with the condition temporary as parameter.

Example for Instrumenting a Spinning Read Loop

Figure 8.9 provides a simple example of an instrumented spinning read loop. The machine code is translated into three superblocks by Valgrind. The first superblock starts at the address `0x400718` for the `jmp` instruction and the second one at the address `0x40071a` for `callq` instruction, which is a function call to a dynamic loaded library. For the sake of simplicity, we only show the third superblock beginning at the address `0x40071f`, where a spinning read loop is detected.

The detected spinning read is located at the address `0x40071f`. The required instrumentation functions `START_SPIN_READING`/`STOP_SPIN_READING` are inserted at the proper place for the related instructions (lines 1 and 5). The memory

```
...
while (spinlock != 0){
  sched_yield();
}
...
```

(a) Spinning read loop.

```
...
400718: jmp   40071f <main+0x34>
40071a: callq 4005a0 <sched_yield@plt>
40071f: mov   0x20091b(%rip),%eax <--
400725: test  %eax,%eax
400727: jne   40071a <main+0x2f>
...
```

(b) Spinning read loop - machine code.

```
-- IMark(0x40071F, 6) --
t5 = LDle:I32(0x601040:
     I64)
t13 = 32Uto64(t5)
t4 = t13
PUT(0) = t4
-- IMark(0x400725, 2) --
t3 = GET:I32(0)
PUT(128) = 0x13:I64
t14 = 32Uto64(t3)
t6 = t14
PUT(136) = t6
PUT(144) = 0x0:I64
-- IMark(0x400727, 2) --
PUT(168) = 0x400727:I64
IR-NoOp
t17 = Shl64(t6,0x20:I8)
t16 = CmpEQ64(t17,0x0:
     I64)
t15 = 1Uto64(t16)
t12 = t15
t18 = 64to1(t12)
t7 = t18
if (t7) goto {Boring} 0
     x400729:I64
goto {Boring} 0x40071A:
     I64
```

(c) Spinning read loop - IR code.

```
DIRTY 1:I1 ::: START_SPIN_READING{0
     x380054b0}()
---- IMark(0x40071F, 6) ----
DIRTY 1:I1 ::: EVH__MEM_HELP_READ[
     rp=1]{0x38010440}(0x601040:I64)
t5 = LDle:I32(0x601040:I64)
t13 = 32Uto64(t5)
t4 = t13
PUT(0) = t4
DIRTY 1:I1 ::: STOP_SPIN_READING{0
     x380054c0}()
---- IMark(0x400725, 2) ----
t3 = GET:I32(0)
PUT(128) = 0x13:I64
t14 = 32Uto64(t3)
t6 = t14
PUT(136) = t6
PUT(144) = 0x0:I64
---- IMark(0x400727, 2) ----
PUT(168) = 0x400727:I64
IR-NoOp
t17 = Shl64(t6,0x20:I8)
t16 = CmpEQ64(t17,0x0:I64)
t15 = 1Uto64(t16)
t12 = t15
t18 = 64to1(t12)
t7 = t18
if (t7) goto {Boring} 0x400729:I64
goto {Boring} 0x40071A:I64
```

(d) Spinning read loop - instrumented IR code.

Figure 8.9.: Example of spinning read loop.

state machine (MSM) traces every read access, whether a spinning read or a normal read operation happens(EVH__MEM_HELP_READ, line 3).

For the case of synchronization with inter-thread event notification, we have to instrument while-loops which enclose a call to the **wait**() library function.

The loop condition is usually not very complex (e.g. a flag or a function call) and the loop body often only contains a call to `wait()` at the beginning. This means that the branching distance is short. We consider branches up to ten basic blocks (approximately equal to ten function calls and branch distances up to 2000 instruction bytes). In the next step, we look inside the loop body and determine whether there is a call to `wait()`. Helgrind$^+$ searches for branches to the `wait()` address in the loop body. An example deals with inter-thread event notification and `wait()` function call which is described in Section 8.5.2.

8.5.1. Look Ahead Instrumentation

Valgrind instruments a program block by block and just-in-time at runtime. Only the piece of code is instrumented that is actually going to be executed. If the program reaches at point that is not instrumented, Valgrind reads a machine code block (superblock) instruments it, and lets the program executes the instrumented code. The instrumentation is done dynamically on the fly.

As before mentioned, the instrumentation itself is divided into three steps: A *disassemble* step, which the machine code is translated into an independent platform intermediate language (IR). The step for *IR instrumentation* that the analyzing tool e.g. Helgrind$^+$ instruments the IR code, and the last step *resynthesize*, where the instrumented IR code is translated back to machine code. Helgrind$^+$ receives the program code piecewise in form of superblocks. It instruments the superblock immediately and gives it back to Valgrind.

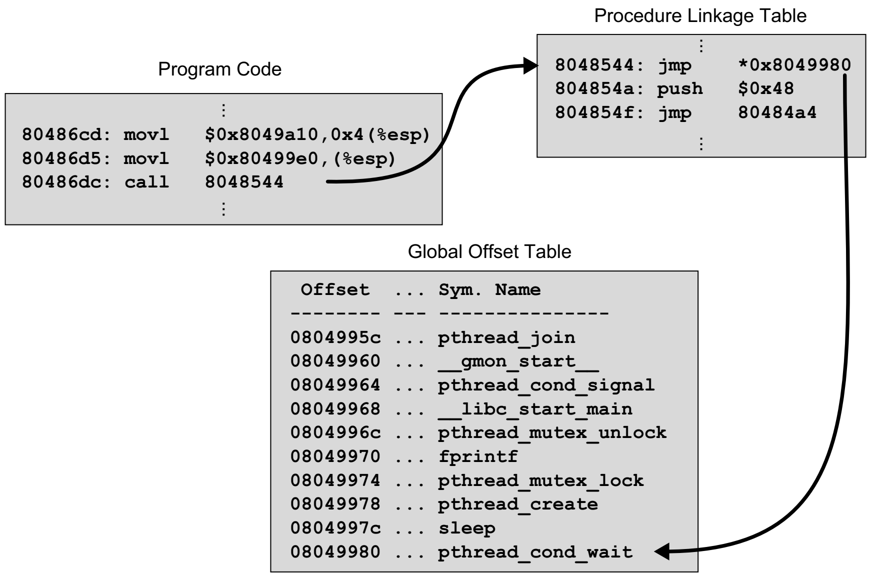
However, during instrumentation, information about a superblock not yet instrumented/executed might be needed. For example, when instrumenting the loop branch, it has to be clear if the loop body contains `wait()` function calls (we would not ignore `wait`-function calls, as discussed in Section 6.6). At the time of instrumenting the branch, the loop body is not yet executed – exactly this information is missing. Another example concerns loop identification: If we consider the executing superblock only, it may not be possible to identify the loops. We may need the information about not yet instrumented/executed superblocks to locate the branch destination. For this reason, we extended Valgrind for *look ahead instrumentation*, so that Helgrind$^+$ can request Valgrind to translate an arbitrary machine code block into IR at any time. So, Helgrind$^+$ is able to analyze not yet instrumented/executed superblocks.

The only prerequisite for the look ahead instrumentation is that the branch addresses in machine code should be known. Unfortunately, this is not always the case. For example, in the x86 architecture, *function pointer calls* need an offset which is specified at runtime and based on this offset the absolute function address is calculated. We mark and instrument these function calls and then, we perform a runtime analysis when the function actually executes. Section 8.5.2.1 discusses how to deal with function pointers and solve this problem.

8.5.2. Identifying Function Calls in IR

After finding the loop boundaries during instrumentation, we search for library function calls within the loop body. We use also look ahead instrumentation for dynamic function calls. Normally, calling a function of a dynamic library requires the following three steps:

- Function parameters and function return address are stored according to the calling convention on the stack or in special registers.
- A function *stub* is called which represents the dynamic library function. This function stub is located in PLT (Procedure Linkage Table) of the *Linux ELF file* [11].
- The function stub determines the *absolute* address of the library function. It stores the address in GOT (Global Offset Table). Future function calls directly jump to the address stored in GOT.



```
                 ⋮
80486cd: movl   $0x8049a10,0x4(%esp)
80486d5: movl   $0x80499e0,(%esp)
80486dc: call   8048544
                 ⋮
```

```
                 ⋮
8048544: jmp    *0x8049980
804854a: push   $0x48
804854f: jmp    80484a4
                 ⋮
```

```
 Offset  ... Sym. Name
-------- --- ---------------
0804995c ... pthread_join
08049960 ... __gmon_start__
08049964 ... pthread_cond_signal
08049968 ... __libc_start_main
0804996c ... pthread_mutex_unlock
08049970 ... fprintf
08049974 ... pthread_mutex_lock
08049978 ... pthread_create
0804997c ... sleep
08049980 ... pthread_cond_wait
```

Figure 8.10.: Calling a dynamic library function.

The entries in GOT table are also known as *jump slots*. In the Linux ELF file, the jump slots and corresponding function names are listed in a table called *relocation table*. We extended Valgrind in a way that is able to read out the relocation table of a Linux ELF file. We implemented this function for *x86* and *amd64* ELF types. Thus, the function call is identified, when a jump slot entry is read and an indirect branch to the value of the read jump slot is done.

Figure 8.11 depicts a complete function call of **pthread_cond_wait**(): lines 1, 2, 5 and 6 calculate the value for the stack pointer. By lines 3, 4 and 7, the processor pushes both parameters for the function call and the return address on the stack. Line 8 loads the value of the jump slot entry. At line 9, the processor jumps on the loaded address. The address of the jump slot entry on stack indicates that the function call is **pthread_cond_wait**(). By calling the dynamic function call **pthread_cond_wait**(), a new superblock starts.

```
1 t6 = GET:I32(16)
2 t5 = Add32(t6,0x4:I32)
3 STle(t5) = 0x8049A10:I32
4 STle(t6) = 0x80499E0:I32
5 t7 = Sub32(t6,0x4:I32)
6 PUT(16) = t7
7 STle(t7) = 0x80486E1:I32
8 t3 = LDle:I32(0x8049980:I32)  <-- address of pthread_cond_wait().
9 goto {Boring} t3
```

Figure 8.11.: Function call of **pthread_cond_wait**() in IR.

In case of inter-thread event notifications, the **wait**() function might be skipped and not executed (lost signals). Thus, ahead instrumentation is applied to locate the **wait**() function call within the loop body. The next step after finding the **wait**() function call is to specify the parameters of **wait**() function. These parameters determine the starting point of hb-relation (the signaling side of a waiting thread). The parameters are only known at runtime during the execution of function. Thus, we simulate the loop body and a portion of the stack/registers to determine these parameters. For instance, in Figure 8.11 between lines 8 and 9 the instrumenting functions for stack simulation will be inserted.

The simulation is done at IR level and does not affect the program semantics at all. During the simulation, all temporaries are renamed so that the arithmetic operations stay safe for the actual execution later on. No store operation is executed and the memory accesses are carefully handled. In this way, we are able to determine the parameters independent of the evaluation of the loop condition even if the whole loop-body is skipped. However, this technique does not work well for few cases which need additional runtime information (e.g. when parameters are on the heap). When we are not able to extract the parameters during the instrumentation, the loop in superblock is instrumented and we complete afterwards the hb-relation at actual runtime based on the data dependency analysis.

8.5.2.1. Function Pointers

Function pointers are used by unpredictable function calls, which are mainly used in C++ (see the keyword `virtual`). They are not easy to handle during instrumentation, because the absolute function address is first known at runtime. Figure 8.12 lists an example of a C++ function pointer and the machine code[4].

```
int test(void){
 return (spin != 0);
}

main() {
...
 int (*test_func) (
      void) = &test;

 while (test_func()){
   sched_yield();
 }
...
}
```

(a) Call to a function pointer.

```
 1 ...
 2 40073a: movq  0x4006eb,-0x8(%rbp)
 3 400741:       00
 4 400742: jmp   400749 <main+0x44>
 5 400744: callq 4005a0 <sched_yield@plt>
 6 400749: mov   -0x8(%rbp),%rax
 7 40074d: callq *%rax      <--
 8 40074f: test  %eax,%eax
 9 400751: jne   400744 <main+0x3f>
10 ...
```

(b) Machine code of the `main()` function.

Figure 8.12.: Example of call to function pointer.

At the address `0x400749` (line 6, the content of a local variable is loaded into the accumulator `%rax`. Then in line 7 at the address `0x40074d`, a function call gets issued and `rax` is interpreted as a function pointer address. Generally, it is not possible to know the address referenced by a function pointer during instrumentation. In the example given, the content of the function pointer will be assigned at runtime by the given address (`0x4006eb`) at line 2. Valgrind creates three different superblocks for this example: A superblock at address `0x40073a`. Another superblock after the function call `callq` beginning from address `0x40074f`; and the function `test()` itself also fits in a new superblock. For the sake of simplicity, we show only the first superblock created at address `0x40073a` in Figure 8.13 (a).

First, our algorithm identifies the loading function pointer at address `0x400749` line 6 in the IR code. Then, it detects the actual spinning read loop happening

[4]Only the machine code of `main()` function is listed.

```
-- IMark(0x40073A, 8) --
t6 = GET:I64(40)
t5 = Add64(t6,0xFFFFFFF8:
     I64)
STle(t5) = 0x4006EB:I64
-- IMark(0x400742, 2) --
-- IMark(0x400749, 4) --
PUT(168) = 0x400749:I64
t7 = Add64(t6,0xFFFFFFF8:
     I64)
t9 = LDle:I64(t7)
PUT(0) = t9
-- IMark(0x40074D, 2) --
PUT(168) = 0x40074D:I64
IR-NoOp
t11 = GET:I64(32)
t10 = Sub64(t11,0x8:I64)
PUT(32) = t10
STle(t10) = 0x40074F:I64
t12 = Sub64(t10,0x80:I64)
goto {Call} t9
```

(a) Call to function pointer (first superblock of the example in Figure 8.12).

```
DIRTY 1:I1 :::
     SET_CURRENT_SB_AND_UC[rp=1]{0
     x380066e0}(0x404B52CB0:I64)
-- IMark(0x40073A, 8) --
t6 = GET:I64(40)
t5 = Add64(t6,0xFFFFFFF8:I64)
DIRTY 1:I1 ::: EVH__MEM_HELP_WRITE
     [rp=1]{0x380106c0}(t5)
STle(t5) = 0x4006EB:I64
-- IMark(0x400742, 2) --
DIRTY 1:I1 ::: START_SPIN_READING
     {0x380054b0}()
-- IMark(0x400749, 4) --
PUT(168) = 0x400749:I64
t7 = Add64(t6,0xFFFFFFF8:I64)
DIRTY 1:I1 ::: EVH__MEM_HELP_READ[
     rp=1]{0x380103a0}(t7)
t9 = LDle:I64(t7)
PUT(0) = t9
DIRTY 1:I1 ::: STOP_SPIN_READING{0
     x380054c0}()
-- IMark(0x40074D, 2) --
PUT(168) = 0x40074D:I64
IR-NoOp
t11 = GET:I64(32)
t10 = Sub64(t11,0x8:I64)
PUT(32) = t10
DIRTY 1:I1 ::: EVH__MEM_HELP_WRITE
     [rp=1]{0x380106c0}(t10)
STle(t10) = 0x40074F:I64
t12 = Sub64(t10,0x80:I64)
goto {Call} t9
```

(b) Call to function pointer - instrumented IR code.

Figure 8.13.: Example of call to function pointer - IR code.

in the `test()` function, since the address of each load operation within the loop is constant.

We instrument each superblock containing a function pointer at the beginning with the analysis function SET_CURRENT_SB_AND_UC as shown in Figure 8.13(b). Therefore, we are able to trace the function pointer and extract the exact address at runtime to handle unpredictable function calls. If any spinning read is detected the functions START_SPIN_READING/STOP_SPIN_READING are inserted as before.

Furthermore, the functions `EVH__MEM_HELP_READ`/`EVH__MEM_HELP_WRITE` are used by the memory state machine (msm) to check whether the executing read/write operation is a spinning read or a counterpart write.

8.6. Data Dependency Analysis

We use runtime data dependency analysis to construct the correct hb-relation between corresponding synchronization parts as described in the previous chapters. After executing each instruction in IR, we have to specify the set of all variables written or read by that instruction and identify data dependencies. In fact after executing the instruction a, we determine the sets $a.use$, the read locations[5] by instruction a, and $a.def$, the written locations by instruction a.

Valgrind translates each machine instruction into a sequence of IR instructions. Every IR instruction is either an arithmetic operation or a load/store operation for a register or memory address. Results and operands are stored in temporaries in each case, as earlier mentioned. Considering the temporaries, we are able to construct the computation tree of each IR instruction and extract the existing data dependencies. Data dependencies for temporaries are stored only for a limited period of time, since temporaries do not exist outside their superblock.

An example is provided in Figure 8.14: On the left side in IR code, lines 3 and 5, two variables are loaded form the stack and then added at line 6. Then, the result is pushed on the stack at line 8. The addresses come from the value of stack pointer at line 1 adding with different offset at lines 2, 4 and 8. On the right side of the figure, the computation tree for the example is depicted. The dashed lines indicate address calculations, whereas data nodes are linked by solid arrows. The numbers in the nodes are the line numbers given by IR code.

The algorithm to obtain the sets of $a.def$ and $a.use$ is simple: Each time when a write command in IR is executed (ST for main memory and PUT for registers), we look in the computation tree for data nodes with a load command (LD for main memory and GET for registers). The locations written by a write command are assigned to $a.def$, whereas the locations read by a load command build $a.use$.

[5] As location is meant a main memory location or a register, identical to a variable as in high level programming languages.

```
t14 = GET:I32(20)
t13 = Add32(t14,0xFFF4:I32)
t15 = LDle:I32(t13)
t12 = Add32(t14,0xFFF8:I32)
t11 = LDle:I32(t12)
t10 = Add32(t11,t15)
t16 = Add32(t14,0xFFF0:I32)
STle(t16) = t10
```

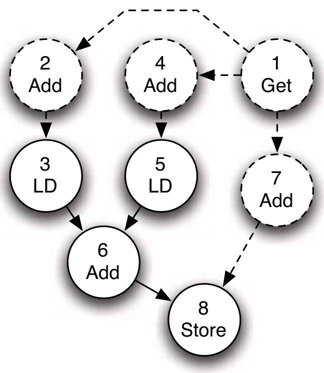


Figure 8.14.: IR code and the corresponding computation tree.

8.7. Limitations

For the implementation, we have considered the library function calls with the calling convention in Linux. For other platforms, we have to adjust our implementation. However, the presented methods and techniques for detecting loops and condition variables are at IR level and independent of a specific platform.

When using nested loops or nested conditions, it may not be easy to find the spinning read loop. Figure 8.1 depicts an example of a nested spinning read loop. The implementation of our algorithm handles such cases as ordinary loops: Each loop is processed as long the loop body is small enough (3-7 superblocks). If the spinning read loop is identified, then it is instrumented.

```
        ...
  /* synchronization */
  while( FLAG1 ) {
      while( FLAG2 ) {
          while( FLAG3 ) {
              /* do_nothing */
          }
      }
   }

   /* do something... */
   do_something();
         ...
```

Listing 8.1: A nested spinning read loop.

If there are many nested loops and loops get bigger than the preset threshold (seven blocks) then the loop might not be identified. Certainly, one could increase the threshold for the number of blocks to be analyzed during loop detection, which might increase the overhead. However, such big nested loops are rarely used as spinning reads.

Listing 8.2 depicts a multi conditions while-loop that *FLAG2* is evaluated only if *FLAG1* is true. The while-loop calls the `wait()` function. In machine code level, such while-loops are converted into nested loops, similar as the above example; they are handled in the same way. Each loop is processed if it is small enough and the end of each loop is considered as the ending synchronization point.

```
while( FLAG1 && FLAG2 ) {
    pthread_cond_wait(...)
}
```

Listing 8.2: Multi-conditions evaluation.

Chapter 9.

Experiments and Evaluation

In this chapter, we present our experiences with Helgrind$^+$ and evaluate our approach by applying it to a number of various applications and benchmarks. Both memory state models, MSM-long and MSM-short, are evaluated. While MSM-short is sensitive and reports immediately a race as soon as it is detected, MSM-long is less sensitive and waits for the reoccurrence of a race in some situations. The new feature for synchronization with inter-thread event notifications is tested. Then, identifying ad-hoc synchronizations in programs is evaluated. Finally, results of Helgrind$^+$ as a universal race detector (neglecting library information) are presented and discussed.

9.1. Experimental Setup

All experiments and measurements in this chapter were conducted on a machine with 2x Intel XEON E5320 Quadcore CPUs at 1.86GHz, 8 GB RAM, running Linux Ubuntu 8.10.1 x64. All programs use POSIX Threads or GNU OpenMP for parallelization. The programs were compiled with GCC 4.2 for x64 architecture. We did not annotate any source code.

Each program was executed 5 times to lessen the effect of random schedules during our measurements. Average values are presented with maximum and minimum values in some cases. For the experiments, we employed different existing detectors such as the commercial tool Intel Thread Checker 3.1 [1] or the open source happens-before race detector *DRD* [47] to compare Helgrind$^+$ against them.

9.2. Results

We applied our approach on various benchmarks suites. The first experiment with Helgrind$^+$ was the analysis of SPLASH-2 benchmarks [43] and several multi-threaded programs, such as a parallel single source shortest path algorithm (sssp) or parallel Bzip, collected from our previous publications [38, 37]. The results show that the algorithm in Helgrind$^+$ is able to reduce the number of false positives without missing races [23]. In this section, we summarize the results of two benchmarks: A unit test suite for race detectors, called *data-race-test* [46], and the recently released *PARSEC* benchmark [2].

9.2.1. Unit Test Suite

We applied Helgrind$^+$ to programs provided in *data-race-test* [46], a unit test suite for race detectors. The test suite aims to create a framework to check a race detector and evaluate the effect of each test case on the tool. It provides more than 150 short programs (called unit test cases) that implement different scenarios which could occur during execution of multi-threaded programs. The scenarios include tricky situations, which are difficult to discover by race detectors. Currently, 120 of these unit test cases can be classified into two main categories: 40 "racy" programs that involve at least one data race, and 80 "race-free" programs. The remaining 30 programs are related to performance and stress testing. We examine and analyze the effect of each test case on Helgrind$^+$. All test cases are implemented in C/C++ using PThreads with a varying number of threads (2-32) and executed without any source code annotation.

9.2.1.1. Evaluation of MSMs and Inter-thread Event Notifications

Table 9.1 shows the results of our experiment. The basic version of Helgrind$^+$ based on MSM-short denoted by *short+lib* in Table 9.2 failed on 41 test cases out of 120. The option *lib* denotes that Helgrind$^+$ uses the library information and intercepts PThreads synchronization primitive calls. It produces 33 false positives and eight false negatives.

When using MSM-long which is less sensitive and optimized for long-running applications, the false positives are reduced to 27 but on the other hand, the false negatives increase to 9. This indicates that MSM-long, denoted by *long+lib* is not adequate to find data races in short programs provided in the unit tests.

Tools	False Positives	False Negatives	FP+FN	Passed Cases
Helgrind$^+$ short+lib	33	8	41	80
Helgrind$^+$ short+lib+cv	28	6	34	86
Helgrind$^+$ long+lib	27	9	36	84
Helgrind$^+$ long+lib+cv	22	9	31	89

Table 9.1.: Results on the test suite *data-race-test*. FP and FN denote False Positives and False Negatives, respectively. *lib* means interception of Pthread library and *cv* enables correct interception of condition variables.

Applying the extended feature for correct handling of inter-thread event notifications increases the accuracy of Helgrind$^+$ for both memory state machines, and reduces the false positives and false negatives. Helgrind$^+$ based on MSM-short with the extended feature denoted by *short+lib+cv* improves the results. Since each unit test is a short program, the results are as expected. Activating the lost signal detection and writ/read dependency analysis removes five false positives and one false negative, i.e., 86 cases of 120 cases pass. Two test cases provided in the unit test suite are similar to the example for spurious wake ups provided in Figure 6.3; they are correctly passed, as we detect the write/read dependency.

All test cases regarding lost signal detection are passed by enabling the *cv* option. Most cases that failed use ad-hoc synchronization defined by the programmer; Helgrind$^+$ is not *yet* aware of them. Few false positives are difficult to identify just by intercepting the synchronization primitives, since they do not follow the standard pattern given for the use of synchronization primitives. Also, few failed test cases are benign data races that should be suppressed.

9.2.1.2. Evaluation of Ad-hoc Synchronizations

We appleid Helgrind$^+$ to the test suite, including checks for ad-hoc synchronization. The results are shown in Table 9.2. We only present the results for Helgrind$^+$ with MSM-short (considering the unit test cases as short-running programs).

The basic version of Helgrind$^+$ (*lib*) in Table 9.2, is only able to intercept PThreada synchronization primitives. It fails on 41 test cases out of 120 and produces 33 false positives with eight false negatives. The new feature *spin(7)* activates spinning read loop detection of up to seven basic blocks. By identifying ad-hoc synchronization, 24 false positives and one false negative are removed, i.e., 106 test cases out of 120 pass. The removed false positives are all apparent data races or synchronization data races that arise from using ad-hoc synchronization. The removed false negative is correctly identified: It was because of spurious wake ups when using the same condition variable by several threads. We consider spinning read loops from three up to seven blocks. All PThreads synchronization primitives are intercepted by Helgrind$^+$.

Tools	**False Positives**	**False Negatives**	**FP+FN**	**Passed Cases**
Helgrind$^+$ lib	33	8	41	79
Helgrind$^+$ lib+spin(7)	8	6	14	106
Helgrind$^+$ nolib+spin(7)	9	6	15	105
Helgrind$^+$ lib+spin(3)	24	6	30	90
Helgrind$^+$ lib+spin(6)	23	6	29	91
Helgrind$^+$ lib+spin(7)	8	6	14	106
Helgrind$^+$ lib+spin(8)	8	6	14	106

Table 9.2.: Results on the test suite *data-race-test*. *lib* means interception of PThreads library; *spin* stands for spinning read detection with the given number of basic blocks as a parameter.

A few failed test cases (false positives) use ad-hoc synchronization. However, it is not easy to detect them, as they do not follow the standard pattern of spinning read loops in the ad-hoc synchronization. They use nested or multi-conditional loops. This is a limitation of our implementation that could be mitigated by improving the algorithm and the implementation.

If we disable PThreads library support, synchronization primitives are no longer intercepted and therefore unknown to Helgrind$^+$. In this case, the detector acts as a pure happens-before detector and no library information is used. We indicate this situation by *nolib+spin(7)* in the table. By this option, only one additional test case fails (one false positive). However, we observe that the best results are achieved when using the new feature as a complementary method to our race detection algorithm (shown as *lib+spin(7)*).

The second part of Table 9.2 depicts the results when using a different number of basic blocks for spinning read loop detection. By increasing the number of basic blocks, the number of false positives are decreased considerably. We got the best result with seven basic blocks. This is because the test suite uses function templates and complex function calls. Thus, spinning read loops in most test cases contain more than three basic blocks. Increasing the number of basic blocks beyond seven did not improve the results further.

9.2.1.3. Comparing with Other Race Detectors

To Compare our results, we employed different race detectors for our measurements: The commercial tool *Intel Thread Checker 3.1* [1] and a 64 bit version of *Helgrind 3.3.1* [45]. Both of them are hybrid race detectors based on lockset and happens-before analyses. Additionally, we compared Helgrind$^+$ against *DRD* [47], an open source happens-before race detector.

The results of Helgrind$^+$ against Helgrind 3.3.1, show significant improvement. 38 more tests passed with Helgrind$^+$ in the best case. This is because of the Eraser-like memory state model used in Helgrind 3.3.1 which performs happens-before analysis for one state only.

Tools	**False Positives**	**False Negatives**	**FP+FN**	**Passed Cases**
Helgrind$^+$ long+lib+cv+spin(7)	7	12	19	101
Helgrind$^+$ long+nolib+spin(7)	6	10	16	104
Helgrind$^+$ short+lib+cv+spin(7)	8	6	14	106
Helgrind$^+$ short+nolib+spin(7)	9	6	16	105
DRD 3.4.1	12	20	32	88
Helgrind 3.3.1	40	12	52	68
Intel Thread checker 3.1	15	21	36	84

Table 9.3.: Results of Helgrind$^+$ and other dynamic race detectors on the test suite *data-race-test*.

We compared the behavior of Helgrind$^+$ with the commercial tool Intel Thread Checker for the test suite. The false positives produced by Intel Thread Checker are nearly twice against false positives produced by Helgrind$^+$. Intel Thread Checker did not detect races in 21 test cases, whereas Helgrind$^+$ overlooked only six racy test cases using the option *short+lib+cv+spin(7)*.

Deactivating PThreads library support (*nolib*), makes Helgrind$^+$ works as a happens-before detector. In addition, we compare our results against the results produced by *DRD 3.4.1* [47], a pure happens-before detector. It achieves considerably better results than DRD. In particular, the number of false negatives with DRD is more than twice than the false negatives with the option *short+nolib+spin(7)*. Having false negatives is a drawback of pure happens-before detectors.

The results on unit tests confirm that Helgrind$^+$ with the provided options is able to discover masked races more accurately compared to other race detectors mentioned here. Especially when using Helgrind$^+$ with the complementary option for the spinning read detection (*short+lib+cv+spin*), the fault detection ratio is promising for small and short-running applications.

9.2.2. PARSEC Benchmark Suite

We applied Helgrind$^+$ to the recently released PARSEC 2.0 benchmark [3]. PARSEC is a benchmark for performance measurement of shared memory computer systems. It differs from other benchmark suites, as it is not HPC-focused. It contains 13 programs from different areas such as computer vision, video encoding, financial analytics, animation, physics and image processing. The programs use different synchronization mechanisms, which makes the benchmark an ideal test case for race detectors. Table 9.4 provides a short summary of the programs[1].

The PARSEC programs support different threading libraries. Most of them use the standard POSIX Thread Library (PThreads) by default. `freqmine` uses OpenMP, and `vips` uses Glib [26] functions for parallelization. Though OpenMP and Glib use internally PThreads for their implementations, their synchronization constructs are unknown to Helgrind$^+$ and not supported. These unknown synchronization constructs causes false positives.

PARSEC offers different input sets varying in size for the program executions. We used the input sets `simsmall` for simulations and ran each program five

[1] `raytrace` was added recently to PARSEC benchmark and comes with the new version. We placed it as the last program in our evaluation order.

Program	**Application Domain**	**Parallelization Model**	**Input Set**	**Thread Model**
blackscholes	Financial Analysis	data-parallel	simsmall	POSIX
bodytrack	Computer Vision	data-parallel	simsmall	POSIX
canneal	Engineering	unstructured	simsmall	POSIX
dedup	Enterprise Storage	pipeline	simsmall	POSIX
facesim	Animation	data-parallel	simsmall	POSIX
ferret	Similarity Search	pipeline	simsmall	POSIX
fluidanimate	Animation	data-parallel	simsmall	POSIX
freqmine	Data Mining	data-parallel	simsmall	OpenMP
streamcluster	Data Mining	data-parallel	simmedium	POSIX
swaptions	Financial Analysis	data-parallel	simmedium	POSIX
vips	Media Processing	data-parallel	simsmall	Glib
x264	Media Processing	pipeline	simsmall	POSIX
raytrace	Visualization	data-parallel	simsmall	POSIX

Table 9.4.: Summary of PARSEC benchmarks.

times, averaging the results. For `streamcluster` and `swaptions`, we use the `simmedium` input set, since the runtime with `simsmall` is too short for our measurements.

Helgrind$^+$ reports data races together with the program context of the second unsynchronized access. If several races in the same program context happen, Helgrind$^+$ reports only the first race context. The numbers provided in the following tables are distinct program code locations with at least one potential data race that we call *racy contexts*.

PARSEC allows the setting to set the number of executing threads. We present the result of executions with two threads for most cases. The empirical study by Lu et al [27] implies that most concurrency bugs manifest themselves with only two threads. Additionally, Valgrind schedules threads in a more fine-grained way than the operating system would do. We assume that many races can already be observed with two threads.

Table 9.5 lists runtime data about PARSEC benchmarks when executing only with two threads. All numbers provided in the table for read/write instructions and synchronization primitives are totals across all threads. Numbers for synchronization primitives include primitives in system libraries. *Locks* are all lock-based synchronizations including *read-write locks* (*rwlocks*). Barriers are barrier-based synchronizations, *Conditions* are waits on condition variables.

The authors of the PARSEC Benchmarks claim the programs to be race free, but

Program	LOC	Instructions (10^9)		Synchronization Primitives		
		Reads	Writes	Lock	Barrier	CV
blackscholes	812	0.092	0.045	0	2	0
bodytrack	10,279	0.425	0.102	35,849	215	90
canneal	4,029	0.435	0.187	88	0	0
dedup	3,689	0.658	0.254	18,436	0	3,536
facesim	29,310	9.632	4.191	10,460	0	1,795
ferret	9,735	0.005	0.002	6,660	0	10
fluidanimate	1,391	0.584	0.144	923,750	0	0
freqmine	2,706	0.744	0.283	78	0	0
streamcluster	1,255	1.795	0.033	146	12,998	34
swaptions	1,494	1.414	0.365	78	0	0
vips	3,228	0.758	0.199	10,575	0	2,698
x264	40,393	0.500	0.204	1,339	0	157
raytrace	13,302	19.260	106.4	867,339	0	3,862

Table 9.5.: Runtime data on PARSEC executed with two threads for input set `simsmall` except `swaptions` and `streamcluster` that are for `simmedium`.

we cannot be absolutely sure od that. Under the assumption that the programs are race free, the warnings produced by race detectors could be counted as false positives. However, we discuss each racy context produced by Helgrind$^+$ and analyze them to see if they are true races or not.

Table 9.6 depicts the results of the experiments with two threads on Helgrind$^+$ with the option for correct condition variables handling (*lib+cv*). As before, *lib* denotes the interception of PThreads (using library information). *short* or *long* symbolizes the memory state machines MSM-short or MSM-long, respectively.

9.2.2.1. Programs without Using Condition Variables

Out of 13 programs in PARSEC, five programs i.e. *blackscholes*, *fluidanimate*, *freqmine*, *swaptions* and *canneal* do not apply any inter-thread event notifications (condition variables). So, the extended option *lib+cv* for correct handling of condition variables should not effect the results. This could be verified in the upper part of Table 9.6. The number of warnings for these five programs are constant whether this option is active or not.

Program	**Helgrind$^+$**		**Helgrind$^+$**	
	short+lib	**short+lib+cv**	**long+lib**	**long+lib+cv**
blackscholes	0	0	0	0
canneal	1	1	1	1
fluidanimate	0	0	0	0
freqmine	149.2	151.4	146.6	149.4
swaptions	0	0	0	0
bodytrack	21.2	15.6	30.2	26.6
dedup	1	0	1	0
facesim	85.3	81	77.4	91.6
ferret	111	2	43.6	2
streamcluster	3.6	2	1	1
vips	48.2	46.6	47.4	46.2
x264	103	28	103	24.8
raytrace	88	82	64	64

Table 9.6.: Number of racy contexts reported on PARSEC with two threads.

blackscholes The program *blackscholes* uses barriers for its synchronization. Every synchronization operation is detected by Helgrind$^+$ and no race is reported.

canneal The threads in *canneal* work also on separated data blocks and do not produce races. However, an intentional data race is reported, on a variable used for ad-hoc synchronizing. The race occurs in the function `annealer_thread::Run()`, which is executed by all threads. All threads accomplish their work in a spinning read loop. Threads check the global variable `_keep_going` as a flag in each iteration to stop or continue to work (spin).

In fact, threads use a programmer-defined synchronization and communicate through the variable with each other in a spinning read loop. The spin variable `_keep_going` is changed only once and is set to `false` to indicate threads to leave the loop. The intentional data race happens as a synchronization data race which should be suppressed. We show in the upcoming results that Helgrind$^+$ is able to remove this benign synchronization race successfully.

fluidanimate *fluidanimate* protects its critical regions with locks. No other synchronization primitives are applied. Helgrind$^+$ produces no warning for this program.

freqmine *freqmine* uses OpenMP to parallelize the program. The Linux implementation of OpenMP uses synchronization operations that are unknown to Helgrind$^+$ and therefore, it cannot detect them. For this reason, many false positives are reported, which increases when using more threads. We show later that spinning read detection suppresses the false positives due to unknown synchronization primitives of OpenMP.

swaptions In *swaptions*, there is only one code block where threads are created. Shortly after the creation, threads are joined. This means the main thread waits till all worker threads finish their jobs, and then continues. The worker threads are all working on separated data blocks. Helgrind$^+$ does not generate any false positive here.

9.2.2.2. Programs Using Condition Variables

The remaining programs in PARSEC use condition variables for synchronization between threads. The lower part of Table 9.6 depicts the results for these programs running with two threads.

bodytrack *bodytrack* uses mainly condition variables. However, *bodytrack* implements its own synchronization functions (**signal**/**wait**) for inter-thread event notifications. Helgrind$^+$ without ad-hoc synchronization detector is not able to identify them. In their implementation, the POSIX variable `cv` is used for sleeping (waiting) and wakening up (signaling) the threads. If a signal is sent in this program, the **signal**`()` primitive of the PThreads library is not called, as in the case of synchronization with a normal condition variable.

Our extension for correct handling of inter-thread event notifications based on direct interception of the PThread primitives does not improve the results. Since PThread primitives are used to implement ad-hoc synchronization (A more abstract synchronization functions at a higher level). Therefore, the results provided in Table 9.6 are as expected: Only minor improvement is achieved because of write/read dependency analysis which makes the parameter `cv` more recognizable. The high level synchronization functions implemented here will be detected by enabling the option of Helgrind$^+$ to detect ad-hoc synchronization, as we show in the following section.

dedup *dedup* utilizes a task queue that divides the work between threads. Each operation on task queue is correctly synchronized by a pair of **signal/wait** primitives that are identified by the option *lib+cv*.

dedup is the only program in PARSEC that applies nested conditions in synchronization with condition variables. Such synchronizations may create multiple ending synchronization points during loop detection (as we discussed in Section 6.6). Apparently, this causes no problem for Helgrind$^+$ and by the enabled option *cv*, all synchronizations by condition variables are correctly identified and no false warning is reported.

facesim This program uses *task queue* as an ad-hoc synchronization to distribute work between threads. Locks and condition variables (signal/wait) of the PThreads library are used to implement the task queue. In *facesim*, on each access to the task queue, synchronization with signal/wait does not necessarily happen.Our race detector can only detect that the task queue is protected by a lock, but does not recognize the programmer-defined synchronization used for data packets delivery through task queue. Thus, the option *lib+cv* does not suppress the false positives because of the inconsistent synchronization use of signal/wait.

ferret *ferret* uses also a task queue, but unlike *facesim*, each access to the task queue is synchronized with condition variables (signal/wait). Before each access, it is checked that the task queue is not empty or is not full. The option *lib+cv* improves the results significantly; only two benign races remain: (1) a variable is used as a flag to signal if the input is read completely, (2) a variable is used as a counter for the input packets. Both variables are modified only once by one thread and are read several times by other threads.

streamcluster In *streamcluster*, we detect five racy contexts by the option *lib*. Three of them are apparent data races counted as false positives. Activating the option *cv* removes these three apparent races (see Table 9.6).

The remaining two racy contexts are benign races: One is an intentional race[2] that no synchronization happens between the initialization and the use of an array in the program. The other benign race is when all threads write the *same* value into a variable[3] using ad-hoc synchronization.

[2] `streamcluster.cpp, Line 202 and Line 236`
[3] `streamcluster.cpp, Line 275`

vips *vips* uses the Glib, a synchronization library unknown to Helgrind$^+$. Additionally, the functions in Glib are called by means of a global variable which contains a collection of function pointers. Therefore, Helgrind$^+$ with *lib+cv* cannot detect the anonymous synchronization operations used in *vips* and generates many false warnings.

x264 Most synchronization operations in *x264* are detected by the option *lib+cv*. However, some ad-hoc synchronizations are still hidden to Helgrind$^+$. Without enabling the option *cv* for inter-thread event notification more than 1000 racy contexts[4] are reported, whereas by activating this option only 28 remain.

By analyzing the warnings, we found benign races on some code blocks. Before the main thread starts a new thread, the whole working context of the old thread is copied by the main thread for the new thread[5]. Each thread uses a part of his working context as a local memory block. The copied data is not used by the new thread, and will be initialized immediately. All reported races are harmless which could be counted as false positives.

raytrace This program uses also a *task queue* as an ad-hoc synchronization to divide tasks between threads. The synchronization primitives from PThreads library i.e. locks and inter-thread event notifications (signal/wait) are used for the implementation of the task queue. Accesses within the task queue, are synchronized by the programmer-defined constructs. Helgrind$^+$ with option *lib+cv* detects that the task queue is protected. However, it cannot identify the programmer-defined synchronizations within the task queue. Therefore, there are still some false positives which Helgrind$^+$ cannot suppress with this option.

In addition to the experiment above, we also provide the results of the same experiment with *four* threads. Table 9.7 lists the results for the option *lib+cv* with *four* threads. Using four threads increases the number of false positives. Generally, by increasing the number of executing threads, the number of apparent data races and synchronization data races will also increase.

Overall, Tables 9.6 and 9.7 have shown that the number of false warnings decreases when using the extended option *cv* for correct handling of inter-thread event notifications.

[4]Helgrind$^+$ reports only the first 1000 racy contexts, the rest will be suppressed.

[5]The copy operation is done within the function `x264_thread_sync_context()`.

Program	Helgrind$^+$		Helgrind$^+$	
	short+lib	short+lib+cv	long+lib	long+lib+cv
blackscholes	0	0	0	0
canneal	1	1	1	1
fluidanimate	0	0	0	0
freqmine	182.6	178.2	170.4	173
swaptions	0	0	0	0
bodytrack	27	23.4	35.8	28.2
dedup	1	0	1	0
facesim	85.3	81	85.3	81.2
ferret	6	2	6	2
streamcluster	5.4	3.6	5.2	3.4
vips	51	50.4	49	47.4
x264	103	49	103	41
raytrace	88	85.4	82	82.4

Table 9.7.: Number of racy contexts reported on PARSEC benchmarks with *four* threads.

9.2.2.3. Ad-hoc Synchronization and Unknown Synchronization Operations

At least eight applications in PARSEC use explicit ad-hoc synchronization in addition to standard synchronization primitives (locks, barriers and condition variables). All programs except `freqmine` and `vips` use PThreads. In this section, we present the results with Helgrind$^+$ based on MSM-short for spinning read loop detection (option *spin*).

The results of our experiments on PARSEC are demonstrated in Table 9.8. Loops identified as a potential spinning read loop by our algorithm, are listed in the third column. Compared to the basic versions of Helgrind$^+$ with option *lib* or *lib+cv*, the number of warnings produced by activating the new feature *spin* is reduced considerably, as expected. In case of `dedup`, `facesim`, `streamcluster`, `vips` and `raytrace`, all false warnings are eliminated. Two benchmarks, `freqmine` and `vips`, use unknown libraries: OpenMP and Glib. The number of warnings decreases to eight and zero respectively. In `x264` and `dedup`, the basic version of Helgrind$^+$ produces more than 1000 warnings (only the first 1000 warnings are reported by the tool), whereas with the new feature (*lib+cv+spin*), all warnings for `dedup` are suppressed and only 32 warnings remain for `x264`. *ferret* generates only two warnings. All together, nine out of 13 applications do not produce any warnings.

We examine the warnings generated by our race detector. All warnings are benign races. The reasons for false warnings in some cases are synchronization constructs such as a *task queue* implemented by the programmer in a different way that does not match the spinning read loop pattern. For instance, consider the program `ferret` that uses a task queue and contains two benign races: A variable is used as a counter for input packets. A single thread modifies it, while other threads read it without any synchronization. Another benign race is a variable that is used as signal to indicate if the input is read completely. In both cases, an obscure implementation is applied and the variables are not used as a condition for a while loop.

Program	**Loops**		**Racy Contexts**		
	Total	**Spins**	**Helgrind$^+$ lib**	**Helgrind$^+$ lib+cv+spin**	**Helgrind$^+$ nolib+spin**
blackscholes	28	8	0	0	0
bodytrack	303.4	79.2	21.2	2	4
canneal	152	43	0	0	0
dedup	116	32	1000	0	0
facesim	1445	171	123.8	0	5
ferret	6301	219	6	2	2
fluidanimate	161	34	0	0	0
freqmine	207	87	149.2	8.4	8.4
streamcluster	100	37	3.6	0	0
swaptions	144	36	0	0	0
vips	583	329	48.2	0	0
x264	383	199	1000	32.2	31.8
raytrace	199.4	77.2	88	0	0

Table 9.8.: Number of potential racy contexts reported on PARSEC benchmarks with two threads.

If we disable PThreads support so that the detector works as a pure happens-before detector based on identification of spinning read loops (*nolib+spin*), approximately the same results are achieved. The number of false positives increases only in two cases. The detector is not aware of the library information and works as a universal race detector.

Overall, the results on various benchmarks confirm that Helgrind$^+$ with the new complementary feature is able to discover ad-hoc synchronization and synchronizations used from unknown libraries without modifying the program or upgrading the race detector. The results are promising and acceptable for real world applications, help programmers to focus on the real data races.

9.2.2.4. Comparison to other Dynamic Race Detectors

To validate our method, we also compared the results of our race detector on PARSEC with other race detectors: a pure happens-before race detector DRD 3.4.1 and two hybrid race detectors: Helgrind 3.1.1 and Intel thread Checker. Table 9.9 compares the results of Helgrind$^+$ for MSM-short against the results of other race detectors.

Program	Helgrind$^+$		DRD	Helgrind	Intel
	lib+cv+spin	nolib+spin	3.4.1	3.3.1	TC
blackscholes	0	0	0	2	0
bodytrack	2	4	31.4	223.6	13
canneal	0	0	0	2	4
dedup	0	0	0	3	0
facesim	0	5	1000	112.6	0
ferret	2	2	246.6	111	0
fluidanimate	0	0	0	58	0
freqmine	8.4	8.4	1000	225.6	1000
streamcluster	0	0	1000	70	2
swaptions	0	0	0	0	0
vips	0	0	838.4	69.4	0
x264	32.2	31.8	1000	486.4	1
raytrace	0	0	1000	117	0

Table 9.9.: Comparing the number of potential racy contexts reported on PARSEC benchmarks for different race detectors. All programs are executed with two threads.

The number of warnings with Helgrind$^+$ are considerably smaller and only few warnings as benign data races are reported. Best results are achieved when using spinning read loop detection plus library interception (option *lib+cv+spin*).

Using spinning read loop detection (*spin*) alone causes Helgrind$^+$ to behave as a pure happens-before race detector. Number of warnings generated in this mode compared to DRD as a happens-before race detector is significantly smaller. DRD produces numerous false alarms – in some cases more than 1000 warnings.

Applying Intel Thread Checker (Intel TC) to PARSEC benchmarks generates acceptable results. However, *lib+cv+spin* in Helgrind$^+$ produces smaller number of warnings in most cases, compared with Intel TC. `vips` uses the `Glib` library that is not supported by Helgrind 3.3.1 and DRD. This is why they show many false alarms. They are not able to intercept synchronization function calls inside `vips`. The same happens in `freqmine` that uses OpenMP. The outcome on PARSEC benchmarks shows that Helgrind$^+$ reports races only in cases where they actually occurred. If any correct synchronization is inferred by Helgrind$^+$, no race is reported. Identifying unknown library calls removes many false positives and makes real world applications easier to be handled by programmers.

Putting all the results together, our empirical results have shown the correctness of our hypotheses (*Hypothesis 1* and *Hypothesis 2*) mentioned in Section 2.3. We have demonstrated that there is a need to define the exact semantics of each synchronization primitive for the detector and have an accurate race detection algorithm. Our experimental data shows clearly that the presented algorithm combines the lockset and happens-before analyses properly, and is able to reduce the number of false positives and false negatives, compared to existing race detectors. So, we prove our first hypothesis (*Hypothesis 1*).

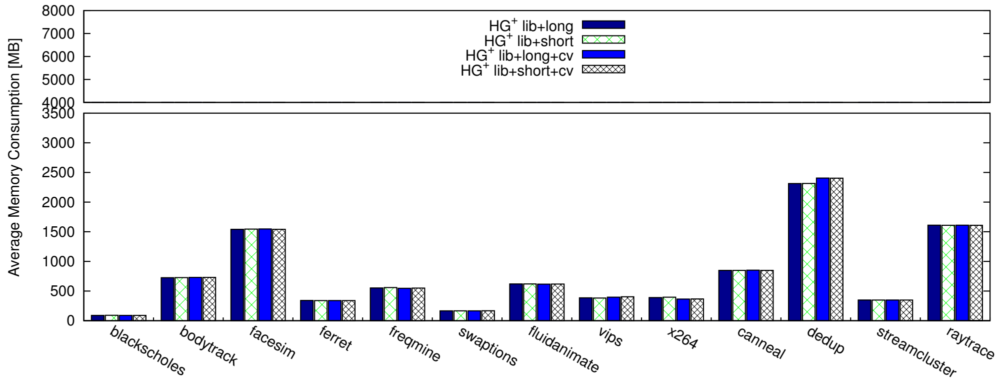
Additionally, we have shown that it is possible to build a universal race detector, which is not limited to a specific set of libraries. Our results confirm that the universal race detector is able to detect different kinds of synchronization operations, and thereby, we have shown the correctness of our second hypothesis (*Hypothesis 2*).

9.3. Performance Evaluation

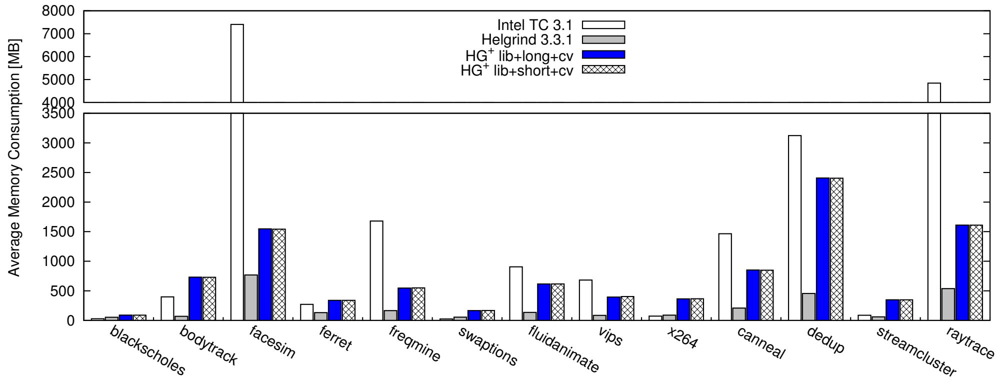
We measured the runtime behavior and the memory requirements of our detector with different features on PARSEC benchmark. The extensions in Helgrind$^+$ cause some reasonable overhead both in time and space. We compared the overhead caused by our detector with other race detectors. All measurements are average values of five executions with two threads using the `simsmall` or `simmedium` inputs for the PARSEC programs. We used `simmedium` inputs for `streamcluster` and `swaptions`, as the runtime with `simsmall` was too short.

9.3.1. Memory Consumption

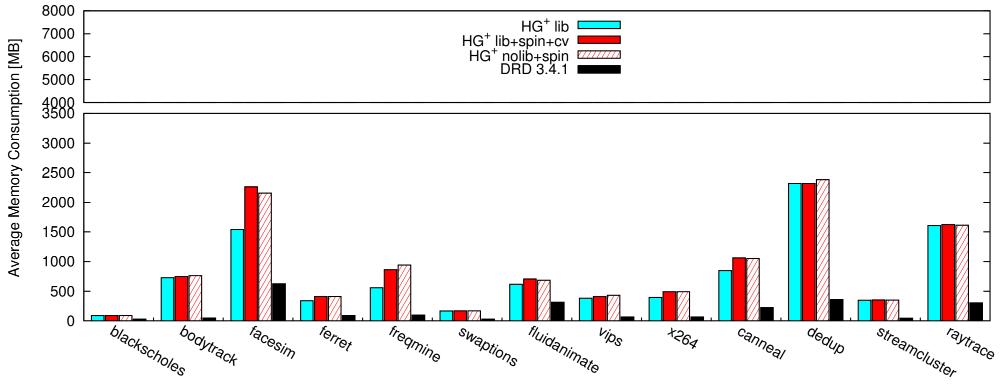
Firstly, we measured the memory usage of instrumented code executed by the detectors. Figure 9.1 depicts the average memory consumption. The memory consumption of Helgrind$^+$ is approximately constant across different memory states machines (Figure 9.1(a)). There is some overhead caused by the extended



(a) Memory consumption for different memory state machines



(b) Memory consumption for handling inter-thread event notifications (option *cv*)



(c) Memory consumption for spinning read loop detection (option *spin*)

Figure 9.1.: Memory consumption on PARSEC by different tools.

memory state machine and the new options implemented features for inter-thread event notifications and ad-hoc synchronizations (options *cv* and *spin*). Some additional memory is required due to the data dependency analysis for deriving the hb-relations.

Figure 9.1(b) compares the overhead of Helgrind$^+$ *lib+cv* with the hybrid detectors. The overhead by Helgrind$^+$ *lib+cv* against Helgrind 3.3.1 is only significant in the case of `dedup`. Although Helgrind 3.3.1 uses a 32-bit shadow memory and Helgrind$^+$ a 64-bit one, the memory overhead by Helgrind$^+$ is not much. The hybrid Intel Thread Checker caused fairly large memory overhead, especially in the case of memory-intensive programs such as `facesim`, `dedup` and `raytrace` that was remarkable while analyzing them with Intel Thread Checker.

Enabling the *spin* option increases memory consumption (Figure 9.1(c)). Since loops in the program have to be checked and analyzed for spinning reads. In case of happens-before detectors as universal race detector: Helgrind$^+$ *nolib+spin* has higher memory than DRD. Note that DRD also uses a 32-bit shadow memory. However, the memory overhead in Helgrind$^+$ is small enough that real world applications with high memory usage are still testable. Optimizing our implementation could help reduce the memory overhead, which we intend to do as a future work.

9.3.2. Runtime Overhead

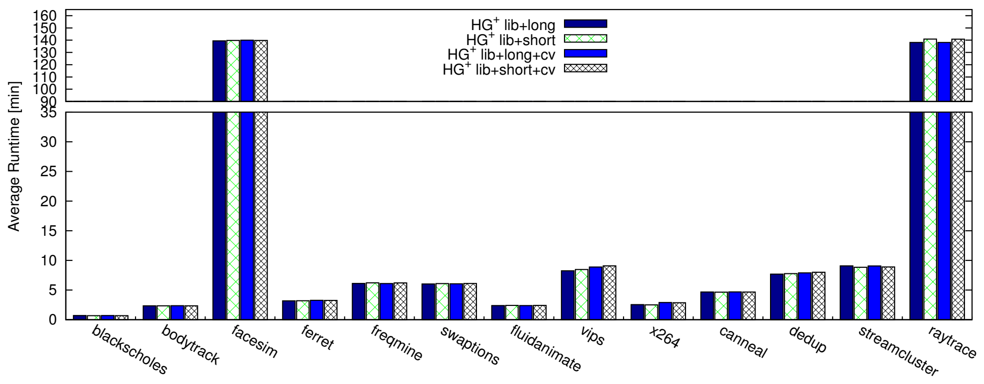
The execution time of instrumented code versus the actual execution time is typically slowed down by a factor of 10 to 70 on Helgrind$^+$. We measured and compared the execution time of instrumented code on different race detectors. The measurements are shown in Figure 9.2 for the PARSEC benchmark. Execution time for different memory state machines are approximately constant (Figure 9.2(a)). The difference in memory state machines incurs almost no runtime overhead.

There is minor overhead of Helgrind$^+$ *lib+cv* over the tool Helgrind 3.3.1 (Figure 9.2(b)). In the worst case, `facesim` and `raytrace` on Helgrind$^+$ increase the execution time significantly. In contrast to other programs, `facesim` and `raytrace` need more time for reading and preparing of the input data. These two programs in PARSEC are object oriented and implemented in *C++*. In other cases, different options of Helgrind$^+$ have approximately equal execution times. Intel Thread Checker increased the execution time remarkably. On average, the slowdown is equal to Helgrind$^+$ *lib+cv*.

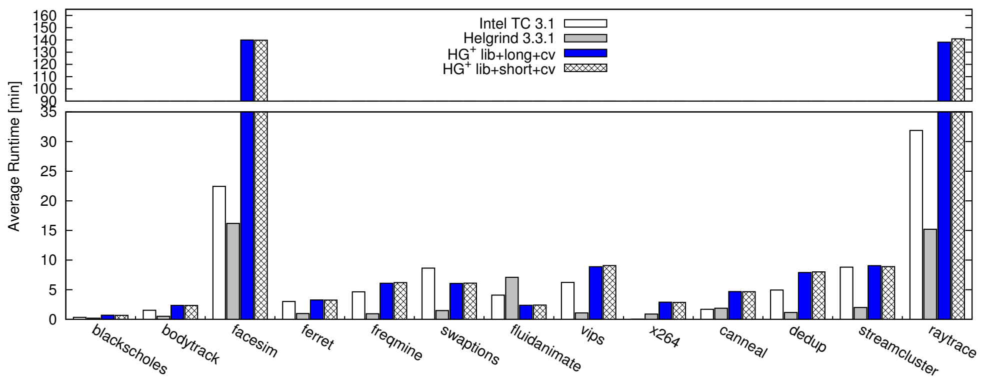
The *spin* option (Figure 9.2(c)) causes some overhead in Helgrind$^+$ compared to the basic mode (*lib* option). In the worst cases, `ferret` and `x264` on Helgrind$^+$ with *lib+spin* and *nolib+spin* options increase the execution time significantly. In other cases, Helgrind$^+$ delivers approximately equal execution times. This is because the main work is done during instrumentation. Normally, code pieces are instrumented only once and used repeatedly during execution. Thus, the slow analysis of loops happens only once. In case of `facesim` and `raytrace`, the execution time by Helgrind$^+$ increases.

Compared to DRD, a happens-before race detector, there is an execution overhead by Helgrind$^+$. But in cases of `dedup` and `fluidanimate` DRD's execution time is much higher. This is because many locks are used in these two benchmarks in contrast to the other programs and DRD has to construct a hb-relation for each lock. On average, the slowdown factor by Helgrind$^+$ with *nolib+spin* as a universal race detector is reasonable to apply for real world applications.

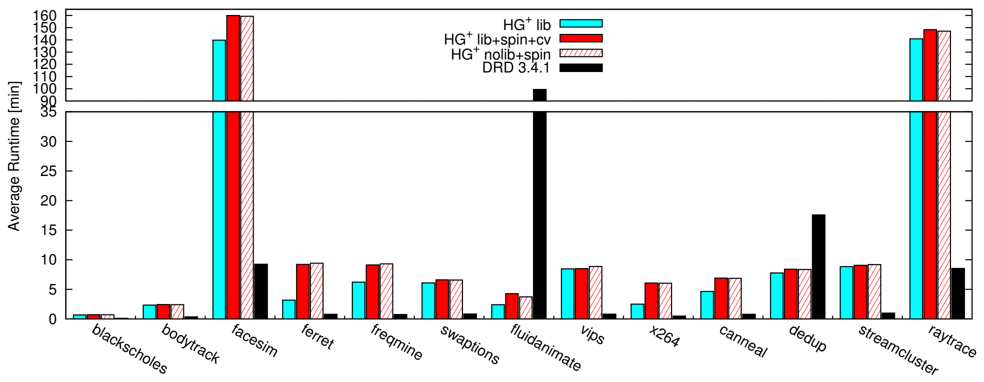
Summing up, the results confirm that the methods presented in this thesis cause reasonable overhead. Helgrind$^+$ is scalable and fast enough and does not need much memory. The true positives are comparable to other tools, and results are accurate in case of short-running as well as long-running applications.



(a) Execution time for different memory state machines



(b) Execution time for handling inter-thread event notifications (option *cv*)



(c) Execution time for spinning read loop detection (option *spin*)

Figure 9.2.: Execution time on PARSEC by different tools.

Chapter 10.

Conclusion

10.1. Conclusion

In this thesis, we presented a novel hybrid dynamic race detection approach based on combining lockset-based and happens-before-based detection. Our approach achieves high accuracy with reasonable overhead by adapting the race detector to short-running and long-running applications.

We extended our hybrid approach to identify synchronization by means of inter-thread event notifications via condition variables. In contrast to classic race detectors, our approach is able to identify synchronization with inter-thread event notifications independent of the execution order. This extension increases the precision of the race detection and reduces the number of false positives and false negatives.

Furthermore, we showed that knowledge of all synchronization operations in a program is crucial for accurate data race detection. We demonstrated that missing ad-hoc synchronization causes numerous false positives. We developed a dynamic method which is able to identify ad-hoc synchronization operations. It has also the ability to detect synchronization primitives from unknown libraries, eliminating the need to modify the detector for each additional library. This method can be used alone or as a complementary method to any other race detector. Using the method alone as a complete race detection approach (with a minor increase in false positives), results in a universal happens-before race detector that is able to detect different synchronization operations from various libraries.

Our hybrid approach is based only on the information extracted by dynamic program analysis. It is automatic and does not need source code annotation or program modifications. The technique that extracts information dynamically by means of just-in-time instrumentation and pre-runtime analysis of the not-yet executed code blocks represents the state-of-the-art.

We implemented our approach in an open source race detector tool called Helgrind$^+$ and applied it to a wide variety of benchmarks. Our empirical results confirm that the new dynamic approach works precisely and removes many false positives and false negatives. The programmer is not overwhelmed by numerous false positives as in other detectors, helping programmers to focus on real races. In situations where certain synchronization patterns occur (e.g. adhoc synchronizations or synchronization with inter-thread event notifications), Helgrind$^+$ is more reliable compared to other race detectors. The evaluation shows that the overhead in our race detector is moderate enough to apply in practical applications.

10.2. Discussion

Our approach reduced the number of false positives and false negatives. However, removing false positives and false negatives in multi-threaded programs completely seems to be impossible. For instance, the program listed in Figure 10.1 demonstrates a false positive that neither Helgrind$^+$ nor any other race detector examined in this thesis could eliminate. The program does not contain any data race, but current race detectors are unable to deal with this type of synchronization. All race detectors used in Chapter 9 report a false data race on `DATA`, whereas the threads are correctly synchronized by any schedule. In fact, Function **try_lock**`(MU)` within `function_2` returns 0 if successful, and accesses on `DATA` are not parallel.

Boehm in [4] shows that a valid interpretation of the Pthread standards requires that a program using the Pthread library must be race free. Furthermore, each access to a shared variable by more than one thread has to be protected by locks. Otherwise, the program behavior is not only non-deterministic but also undefined. Sometimes, it is not easy to discover the programmer's intention and the behavior of programs just by code inspection. Additionally, compilers and processors complicate this task by reordering the instructions and code optimization.

As an example, assume the program provided in Figure 10.2 that initializes variables `X, Y, t1, t2` to zero. Each thread assigns a value to pairs `(X, t1)` or `(Y, t2)`. This causes data races on the variables `X` and `Y`. After executing each code block, `t1` or `t2` would be set to one, depending on whether `X = 1` or `Y = 1` was executed. If the compiler reorders the instructions, it is also possible that `t1 = Y` or `t2 = X` run first. Thus, the values of `t1, t2` are not determined – the behavior of this program is non-deterministic and also undefined.

```
int DATA = 0;

main()
{
  create(thread_1, function_1);
  create(thread_2, function_2);
}
```

```
void function_1() {
  DATA = 1;

  lock(MU);
  return NULL;
}
```

(a) Thread 1

```
void function_2() {
  while (try_lock(MU) == 0)
    unlock (MU);

  DATA = 2;
}
```

(b) Thread 2

Figure 10.1.: A race free program, but detectors report a false data race on DATA.

To resolve this problem, one may put the instructions in a critical section by using a common lock. This solves any atomicity violations if they exist, and eliminates the race warnings caused by them. But the problem still exists and locks cannot help. The program remains non-deterministic and this is because of an *order violation* which happens between threads. In fact, order violations are problems and they are still difficult to guess by detectors – the correct interleaving and the programmer's intention can not be identified completely by race detectors (including Helgrind$^+$).

10.3. Future Research

There are many opportunities for future research to improve Helgrind$^+$. For example, program analysis might be used to automatically select the appropriate memory state machine for a program. Classification techniques for warnings might draw upon the state machine's history. Applying a runtime analysis that excludes variables that are only accessed by a single thread could improve performance as well. Another direction for future work is improving the accuracy of the universal race detector by identifying lock operations, in order to enable lockset analysis.

So far, we have focused on races that might happen only on a single variable. A method to detect races on correlated variables in programs could be very useful.

```
int X = 0;
int Y = 0;
int t1 = 0;
int t2 = 0;

main()
{
  create(thread_1, function_1);
  create(thread_2, function_2);
}
```

```
function_1() {
  X = 1;
  t1 = Y;
}
```

(a) Thread 1

```
function_2() {
  Y = 1;
  t2 = X;
}
```

(b) Thread 2

Figure 10.2.: A program demonstrates an order violation. Even if each function is protected by a **lock**/**unlock** pair, the problem still remains.

Automatically finding the atomic regions and correlated variables in a program at runtime is challenging.

Finally, performing a static or dynamic analysis to reduce the amount of instrumentation improves performance. Since many reads and writes are not parallel, finding a method to exclude them during instrumentation could be quite useful.

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Appendix A.

Helgrind$^+$ User Manual

Helgrind$^+$ 3.4.1 is based on Valgrind 3.4.1, a dynamic binary instrumentation framework. In addition to the standard commands provided by Valgrind into Helgrind$^+$, we shortly explain the new command for the features implemented in Helgrind$^+$. For more details please use the help command provided by the tool.

A.1. Choosing a Memory State Machine

Commands provided in Helgrind+ for choosing the proper memory state machine (MSMs) is as following:

- valgrind –tool=helgrind-ukas: using Helgrind$^+$ with MSM-short (adequate for short-running applications).

- valgrind –tool=helgrind-ukal : using Helgrind$^+$ with MSM-long (adequate for long-running applications).

- valgrind –tool=helgrind-ukas-nolib : using Helgrind$^+$ with MSM-short and turning off the PThreads library interception (no library information).

- valgrind –tool=helgrind-ukal-nolib : using Helgrind$^+$ with MSM-long version and turning off the PThreads library interception (no library information).

A.2. Handling Inter-thread Event Notifications

Command line options for correct handling of inter-thread event notifications (Detecting lost signals and spurious wake ups):

- –lsd=no—yes—wr (could be used for all versions of MSMs).

  –lsd=no : no ahead instrumentation.

  –lsd=yes : use ahead instrumentation for lost signal detection.

  –lsd=wr : with data dependency analysis (constructing write/read-relation) additionally to ahead translation (most accurate option).

- –verbose-lostsig : verbose option to dump details used for debugging purposes.

A.2.1. Example

Running Helgrind$^+$ based on MSM-short with lost signal detection and data dependency analysis (write/read - relation)

- valgrind –tool=helgrind-ukas –lsd=wr –suppressions=helgrind.supp date

The option –suppressions=helgrind.supp could provide a suppression file (if needed) for Helgrind$^+$.

A.3. Spinning Read Loop Detetcion

Option for control flow analysis for spinning read loop detection:

- use ”–cfg=N” to set the maximum number of basic blocks a spinning read loop may span.
  - if it’s set too low, spin reads may be missed and if it’s set too high, the overhead may increase.
  - default value of 3 shows good results for some x86 programs.
- use command line option ”–verbose-cfg” to dump details of detected spin reads.

- use –ignore-pthread-spins: Option to turn off spinning read loop detection within PThreads primitives. This option is only applicable when library primitives are intercepted and it has no influence when *nolib* (e.g. –tool=helgrind-ukas-nolib) is used.

A.3.1. Example

Running Helgrind$^+$ based on MSM-short with spinning read loop detection. Loops with maximum number of basic blocks three is considered for spinning red detection.

- valgrind –tool=helgrind-ukas-nolib –cfg=3 date

In the esxample below, Helgrind$^+$ works as a universal race detector with no interception of library information. Number of basic blocks is set to seven.

- valgrind –tool=helgrind-ukas-nolib –cfg=4 date

A.3.2. Control flow graph

Show the graph output in format dot (graphviz):

- –sow-cfg : output control flow graph.

A.4. Miscellaneous Commands

Some other useful commands implemented for Helgrind$^+$ are listed below:

- -v : Counting the number of calls to synchronization primitives(pthead_mutex, pthread_cond_wait/signal and pthread_barrier_wait).
- –xml=yes—no : XML Output for Helgrind$^+$.
- –hgdebug=useOrigAlgo : Change spin read detection algorithm to the initial algorithm for small loops (for debugging purpose). By default the advance algorithm explained in the thesis is used for spinning read detection.

Appendix B.

Experiment Results on Unit Test Suite

B.1. Results Based on MSM-short

In the following section, we present the detailed results of our race detectors on unit test suite data-race-test[46]. Different options used for Helgrind$^+$ are listed below:

- Helgrind$^+$ +lib: Interception of synchronization primitives from PThreads library.
- Helgrind$^+$ +lib+cv: Interception of PThreads library and correct handling of inter-thread event notification via condition variables.
- Helgrind$^+$ +lib+cv+spin(n): In addition to the library interception and handling of event notifications, Helgrind$^+$ uses spinning read loop detection. n denotes the maximum number of basic blocks during loop detection.
- Helgrind$^+$ +nolib+spin(n): No library interception. Detector works only based on spinning read loop detection and as a pure universal race detector.

Furthermore, we present the detailed results on three other race detectors used for our experiments: DRD 3.4.2 [47], Helgrind 3.3.1 [45] and Intel Thread checker 3.1 [1]. We firstly show the results based on MSM-short and then the results produced by MSM-long.

Test Unit Number	Expected Result	DRD 3.4.1		Helgrind 3.3.1		Intel Thread Checker 3.1		Helgrind+ lib		Helgrind+ lib+cv	
0	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
1	positive	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
2	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
3	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
4	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
5	negative	negative	TN	positive	FP	negative	TN	positive	FP	negative	TN
6	negative	negative	TN	positive	FP	negative	TN	positive	FP	negative	TN
7	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
8	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
9	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
10	positive	negative	FN	negative	FN	negative	FN	positive	TP	positive	TP
11	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
12	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
13	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
14	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
15	negative	negative	TN	negative	TN	negative	TN	positive	FP	positive	FP
16	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
17	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
18	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
19	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
20	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
21	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
22	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
23	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
24	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
25	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
26	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
27	negative	negative	TN	positive	FP	positive	FP	positive	FP	positive	FP
28	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
29	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
30	negative	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
31	negative	positive	FP	positive	FP	negative	TN	positive	FP	positive	FP
32	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
33											
34											
35											
36	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
37	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
38	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
39	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
40	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
41	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
42	negative	negative	TN	positive	FP	negative	TN	positive	FP	negative	TN
43	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
44	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
45	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
46	positive	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN
47	positive	negative	FN	positive	TP	negative	FN	positive	TP	positive	TP
48	positive	negative	FN	negative	FN	positive	TP	positive	TP	positive	TP
49	positive	negative	FN	negative	FN	positive	TP	positive	TP	positive	TP

Test Unit Number	Expected Result	DRD 3.4.1		Helgrind 3.3.1		Intel Thread Checker 3.1		Helgrind+ lib		Helgrind+ lib+cv	
50	positive	positive	TP	negative	FN	negative	FN	negative	FN	negative	FN
51	positive	negative	FN	positive	TP	negative	FN	positive	TP	positive	TP
52	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
53	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
54	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
55											
56	positive	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
57	negative	negative	TN	negative	TN	positive	FP	negative	TN	negative	TN
58	negative	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
59	negative	positive	FP	positive	FP	negative	TN	positive	FP	positive	FP
60	negative	negative	TN	negative	TN	negative	TN	positive	FP	negative	TN
61	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
62											
63											
64	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
65	positive	negative	FN	positive	TP	negative	FN	negative	FN	negative	FN
66	negative	negative	TN	positive	FP	negative	TN	positive	FP	negative	TN
67	positive	negative	FN	positive	TP	negative	FN	negative	FN	positive	TP
68	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
69	negative	positive	FP	positive	FP	negative	TN	positive	FP	positive	FP
70	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
71	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
72											
73											
74											
75	negative	negative	TN	negative	TN	positive	FP	negative	TN	negative	TN
76	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
77	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
78	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
79	negative	negative	TN	negative	TN	negative	TN	positive	FP	positive	FP
80	negative	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
81	negative	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
82											
83	positive	positive	TP	negative	FN	positive	TP	positive	TP	positive	TP
84	positive	positive	TP	negative	FN	positive	TP	positive	TP	positive	TP
85											
86	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
87	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
88	negative	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
89											
90	negative	negative	TN	negative	TN	negative	TN	positive	FP	positive	FP
91	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
92	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
93											
94	positive	negative	FN	positive	TP	negative	FN	positive	TP	positive	TP
95	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
96	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
97	positive	positive	TP	negative	FN	negative	FN	positive	TP	positive	TP
98	negative	negative	TN	positive	FP	positive	FP	positive	FP	positive	FP
99											
100											

Test Unit Number	Expected Result	DRD 3.4.1		Helgrind 3.3.1		Intel Thread Checker 3.1		Helgrind+ lib		Helgrind+ lib+cv	
101	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
102	positive	positive	TP	negative	FN	negative	FN	positive	TP	positive	TP
103	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
104	positive	positive	TP	positive	TP	negative	FN	positive	TP	positive	TP
105	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
106	negative	negative	TN	negative	TN	negative	TN	positive	FP	positive	FP
107											
108	negative	negative	TN	negative	TN	positive	FP	positive	FP	positive	FP
109	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
110	positive	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
111	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
112											
113											
114	negative	positive	FP	negative	TN	positive	FP	negative	TN	negative	TN
115	negative	positive	FP	positive	FP	negative	TN	positive	FP	positive	FP
116	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
117	negative	positive	FP	negative	TN	positive	FP	negative	TN	negative	TN
118											
119	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
120	positive	negative	FN	negative	FN	positive	TP	positive	TP	positive	TP
121	positive	positive	TP	negative	FN	negative	FN	negative	FN	negative	FN
122	positive	negative	FN	positive	TP	negative	FN	positive	TP	positive	TP
123	positive	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
124											
125	negative	positive	FP	positive	FP	negative	TN	negative	TN	negative	TN
126	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
127											
128											
129	negative	negative	TN	positive	FP	negative	TN	negative	TN	negative	TN
130	negative	negative	TN	negative	TN	positive	FP	negative	TN	negative	TN
131	negative	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
132	positive	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
133	positive	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
134	negative	negative	TN	negative	TN	negative	TN	positive	FP	positive	FP
135	negative	negative	TN	negative	TN	positive	FP	negative	TN	negative	TN
136											
137	positive	negative	FN	positive	TP	positive	TP	positive	TP	positive	TP
138	positive	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN
139	positive	negative	FN	positive	TP	negative	FN	positive	TP	positive	TP
140	negative	negative	TN	positive	FP	negative	TN	positive	FP	positive	FP
141	negative	negative	TN	positive	FP	positive	FP	positive	FP	positive	FP
142											
143											
144											
145											
146	positive	positive	TP	positive	TP	positive	TP	negative	FN	negative	FN

Test Unit Number	Expected Result	DRD 3.4.1	Helgrind 3.3.1	Intel Thread Checker 3.1	Helgrind+ lib	Helgrind+ lib+cv
Total						
FP		**12**	**40**	**15**	**33**	**28**
FN		**20**	**12**	**21**	**7**	**6**
FP+FN		**32**	**52**	**36**	**40**	**34**
TP	**40**	**20**	**28**	**19**	**33**	**34**
TN	**80**	**68**	**40**	**65**	**47**	**52**

B.2. Result Based on MSM-long

As previous section, all experiments are done also on MSM-long. Same options of Helgrind$^+$ are used with comparison to other data race detectors. Detailed results are listed as following.

Test Unit Number	Helgrind+ lib+cv+spin3		Helgrind+ lib+cv+spin6		Helgrind+ lib+cv+spin7		Helgrind+ lib+cv+spin8		Helgrind+ nolib+spin7	
0	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
1	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
2	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
3	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
4	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
5	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
6	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
7	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
8	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
9	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
10	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
11	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
12	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
13	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
14	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
15	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
16	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
17	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
18	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
19	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
20	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
21	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
22	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
23	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
24	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
25	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
26	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
27	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
28	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
29	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
30	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
31	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
32	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
33										
34										
35										
36	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
37	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
38	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
39	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
40	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
41	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
42	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
43	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
44	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
45	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
46	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN
47	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
48	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
49	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP

Test Unit Number	Helgrind+ lib+cv+spin3		Helgrind+ lib+cv+spin6		Helgrind+ lib+cv+spin7		Helgrind+ lib+cv+spin8		Helgrind+ nolib+spin7	
50	negative	FN	negative	FN	negative	FN	negative	FN	positive	TP
51	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
52	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
53	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
54	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
55										
56	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
57	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
58	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
59	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
60	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
61	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
62										
63										
64	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
65	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN
66	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
67	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
68	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
69	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
70	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
71	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
72										
73										
74										
75	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
76	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
77	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
78	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
79	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
80	positive	FP	positive	FP	negative	TN	negative	TN	positive	FP
81	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
82										
83	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
84	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
85										
86	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
87	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
88	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
89										
90	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
91	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
92	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
93										
94	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
95	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
96	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
97	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
98	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
99										
100										

Test Unit Number	Helgrind+ lib+cv+spin3		Helgrind+ lib+cv+spin6		Helgrind+ lib+cv+spin7		Helgrind+ lib+cv+spin8		Helgrind+ nolib+spin7	
101	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
102	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
103	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
104	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
105	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
106	positive	FP	negative	TN	negative	TN	negative	TN	positive	FP
107										
108	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
109	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
110	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
111	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
112										
113										
114	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
115	positive	FP	negative	TN	negative	TN	negative	TN	negative	TN
116	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
117	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
118										
119	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
120	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
121	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN
122	positive	TP	positive	TP	positive	TP	positive	TP	negative	FN
123	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
124										
125	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
126	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
127										
128										
129	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
130	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
131	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
132	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
133	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
134	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
135	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
136										
137	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
138	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN
139	positive	TP	positive	TP	positive	TP	positive	TP	positive	TP
140	negative	TN	negative	TN	negative	TN	negative	TN	negative	TN
141	positive	FP	positive	FP	positive	FP	positive	FP	positive	FP
142										
143										
144										
145										
146	negative	FN	negative	FN	negative	FN	negative	FN	negative	FN

Test Unit Number	Helgrind+ lib+cv+spin3	Helgrind+ lib+cv+spin6	Helgrind+ lib+cv+spin7	Helgrind+ lib+cv+spin8	Helgrind+ nolib+spin7
Total					
FP	**10**	**8**	**7**	**7**	**9**
FN	**6**	**6**	**6**	**6**	**6**
FP+FN	**16**	**14**	**13**	**13**	**15**
TP	**34**	**34**	**34**	**34**	**34**
TN	**70**	**72**	**73**	**73**	**71**

Appendix C.

Experiment Results on PARSEC Benchmark

Detailed results of our experiments on PARSEC 2.0 [3]with Helgrind$^+$ and different tools are listed below. All the programs are executed with two threads and the result are the average value of five executions. We used the input sets `simsmall` for simulations. Only, `streamcluster` and `swaptions`, use the `simmedium` input set.

For Helgrind$^+$, we use the following option:

- Helgrind$^+$ +short+lib: Interception of synchronization primitives from PThreads library based on MSM-short.
- Helgrind$^+$ +long+lib: Interception of synchronization primitives from PThreads library based on MSM-long.
- Helgrind$^+$ +short(long)+lib+cv: Interception of PThreads library using MSM-short (MSM-long) and correct handling of inter-thread event notification via condition variables.
- Helgrind$^+$ +lib+cv+spin(n): In addition to the library interception and handling of event notifications, Helgrind$^+$ uses spinning read loop detection. n denotes the maximum number of basic blocks during loop detection.
- Helgrind$^+$ +nolib+spin(n): No library interception. Detector works only based on spinning read loop detection and as a pure universal race detector.

As before, we present the detailed results on three other race detectors used for our experiments: DRD 3.4.2, Helgrind 3.3.1 and Intel Thread checker 3.1.

Helgrind+ long+lib

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	90902.4	87100
bodytrack	30.2	18	39	742201.6	708020
facesim	77.4	69	96	1577336.8	1572588
ferret	43.6	40	46	348602.4	343492
freqmine	146.6	139	152	564195.2	524060
swaptions	0	0	0	169924.8	169008
fluidanimate	0	0	0	634580.8	630060
vips	47.4	45	54	395264.8	387256
x264	1000	1000	1000	399023.2	396936
canneal	0	0	0	868784.8	862860
dedup	1000	1000	1000	2368116.8	2366944
streamcluster	1	1	1	356363.2	355248
raytrace	64	64	64	1649264	1641584

Helgrind+ long+lib

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	92308	41.264	41.23	41.29
bodytrack	752276	140.378	140.10	140.81
facesim	1581880	8367.872	8303.98	8433.36
ferret	354768	190.234	189.11	191.05
freqmine	590892	365.96	365.42	367.45
swaptions	170900	361.262	361.10	361.46
fluidanimate	636184	142.072	141.68	142.45
vips	402924	495.43	491.63	499.93
x264	400576	152.352	151.17	153.79
canneal	875808	279.578	279.52	279.68
dedup	2369956	461.03	459.13	462.47
streamcluster	357464	544.148	540.50	547.96
raytrace	1655168	8288.014	8282.74	8292.43

Helgrind+ long+lib+cv

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	91254.4	88256
bodytrack	26.6	12	39	748012	733224
facesim	91.6	72	128	1583560	1582336
ferret	2	2	2	347500	343348
freqmine	149.4	142	153	559017.6	521448
swaptions	0	0	0	170421.6	169180
fluidanimate	0	0	0	630971.2	630144
vips	6	6	6	405184	395420
x264	24.8	24	25	372440	369104
canneal	0	0	0	872732.8	866688
dedup	0	0	0	2463892	2461664
streamcluster	1	1	1	356170.4	355700
raytrace	64	64	64	1648989.6	1644244

Helgrind+ long+lib+cv

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	92568	41.738	41.71	41.77
bodytrack	761796	141.432	140.90	141.84
facesim	1584464	8394.402	8324.50	8466.55
ferret	350000	196.166	194.37	197.86
freqmine	585308	365.666	364.90	366.37
swaptions	172676	363.124	362.77	363.41
fluidanimate	631832	142.172	141.83	142.44
vips	411800	532.41	527.49	536.98
x264	374920	173.728	167.03	176.19
canneal	881180	280.386	280.25	280.54
dedup	2465352	473.886	472.53	475.00
streamcluster	356680	543.214	542.28	545.44
raytrace	1654840	8286.474	8280.62	8295.30

Helgrind+ short+lib

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	93006.4	90144
bodytrack	21.2	19	23	744834.4	721256
facesim	123.8	109	139	1581156	1578492
ferret	111	111	111	346757.6	343288
freqmine	149.2	140	159	571577.6	545116
swaptions	0	0	0	169684.8	169224
fluidanimate	0	0	0	633720	629916
vips	48.2	41	52	391052.8	388972
x264	1000	1000	1000	405447.2	397964
canneal	0	0	0	868282.4	863972
dedup	1000	1000	1000	2370535.2	2368512
streamcluster	3.6	2	4	357374.4	355164
raytrace	117	117	117	1647012	1642848

Helgrind+ short+lib

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	99532	40.80	40.77	40.82
bodytrack	759364	140.512	139.60	141.38
facesim	1582764	8384.178	8281.12	8516.44
ferret	351728	191.296	190.21	192.06
freqmine	599232	372.532	370.33	373.77
swaptions	169876	365.688	365.31	366.03
fluidanimate	636384	144.98	144.58	145.39
vips	394664	508.132	502.89	511.66
x264	411344	150.368	149.75	150.95
canneal	871136	279.082	278.97	279.30
dedup	2371696	466.368	465.37	467.96
streamcluster	364860	530.242	529.82	530.77
raytrace	1650640	8454.308	8448.77	8460.30

Helgrind+ short+lib+cv

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	91284.8	87608
bodytrack	27.4	19	38	747292.8	737448
facesim	111.2	93	127	1578487.2	1574036
ferret	2	2	2	347194.4	343832
freqmine	151.4	144	155	563044.8	506844
swaptions	0	0	0	171621.6	169952
fluidanimate	0	0	0	631828	630520
vips	6.4	6	8	413779.2	404524
x264	28	28	28	374814.4	371052
canneal	0	0	0	868695.2	863488
dedup	0	0	0	2461259.2	2460336
streamcluster	2	2	2	356942.4	355180
raytrace	85.2	64	117	1647758.4	1643644

Helgrind+ short+lib+cv

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	92808	41.182	41.15	41.21
bodytrack	759804	140.686	139.82	141.19
facesim	1582212	8384.632	8263.11	8532.74
ferret	354084	195.308	194.78	195.98
freqmine	589400	372.15	371.43	372.99
swaptions	177852	366.62	366.46	366.94
fluidanimate	635576	144.906	144.33	145.54
vips	420960	544.326	543.00	545.41
x264	377340	171.048	166.49	174.51
canneal	878932	279.684	279.48	279.89
dedup	2462144	480.072	477.25	482.65
streamcluster	358780	533.674	532.38	536.02
raytrace	1653876	8449.732	8444.02	8457.14

Helgrind+ short+lib+cv+spin3

Package	**Contexts (avg)**	**Contexts (min)**	**Contexts (max)**	**Memory (avg)**	**Memory (min)**
blackscholes	0	0	0	93013.6	89916
bodytrack	0	0	0	769131.2	742332
facesim	0	0	0	2313668.8	2310944
ferret	2	2	2	422665.6	410600
freqmine	8.4	7	10	883900	837284
swaptions	0	0	0	172425.6	171360
fluidanimate	0	0	0	722147.2	719060
vips	0	0	0	421679.2	408780
x264	32.2	29	34	502725.6	498516
canneal	0	0	0	1086510.4	975280
dedup	0	0	0	2370365.6	2369548
streamcluster	0	0	0	360940	358636
raytrace	0	0	0	1665812	1647540

Helgrind+ short+lib+cv+spin3

Package	**Memory (max)**	**Seconds (avg)**	**Seconds (min)**	**Seconds (max)**
blackscholes	95288	42.916	42.79	42.97
bodytrack	787948	145.672	145.14	146.33
facesim	2318164	9591.378	9577.17	9614.52
ferret	454440	552.626	530.28	564.64
freqmine	916936	547.356	529.29	588.06
swaptions	173280	397.066	395.68	398.15
fluidanimate	728640	256.172	255.88	256.42
vips	436396	510.336	508.16	513.95
x264	509312	364.038	360.91	367.88
canneal	1525600	414.128	409.61	429.05
dedup	2371840	504.154	502.45	506.71
streamcluster	363932	542.992	530.19	547.40
raytrace	1716016	8900.76	8891.79	8915.75

Helgrind+ short+nolib+spin3

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	93269.6	92452
bodytrack	12	12	12	781828	749480
facesim	5	5	5	2207486.4	2177052
ferret	2	2	2	421828	401748
freqmine	8.4	7	9	963451.2	844220
swaptions	0	0	0	172263.2	171676
fluidanimate	0	0	0	702328.8	701604
vips	1	1	1	442876	408012
x264	31.8	30	33	503490.4	502016
canneal	0	0	0	1079532	967920
dedup	0	0	0	2436667.2	2435412
streamcluster	0	0	0	358192	356620
raytrace	0	0	0	1653135.2	1646300

Helgrind+ short+nolib+spin3

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	93880	42.728	42.63	42.79
bodytrack	801020	144.774	144.35	145.03
facesim	2323728	9557.668	9520.34	9649.75
ferret	446620	564.93	526.90	629.83
freqmine	1047560	558.058	527.59	575.45
swaptions	173316	394.874	393.83	395.56
fluidanimate	703344	224.278	223.75	224.71
vips	464136	531.432	506.98	548.37
x264	505364	361.95	359.59	365.31
canneal	1514712	411.482	407.62	426.02
dedup	2437656	501.30	499.89	502.93
streamcluster	360352	551.266	543.27	564.71
raytrace	1658128	8828.056	8822.07	8838.03

DRD 3.4.1

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	30050.4	28708
bodytrack	31.4	25	41	47868	46852
facesim	1000	1000	1000	638545.6	636756
ferret	246.6	17	1000	92559.2	87948
freqmine	1000	1000	1000	98746.4	97168
swaptions	0	0	0	29768	29488
fluidanimate	0	0	0	321482.4	318468
vips	838.4	816	855	66500.8	65928
x264	1000	1000	1000	65944.8	65124
canneal	0	0	0	231287.2	230988
dedup	0	0	0	369844	367304
streamcluster	1000	1000	1000	45520	45196
raytrace	1000	1000	1000	309837.6	309056

DRD 3.4.1

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	33928	6.926	6.89	6.99
bodytrack	48836	21.534	21.52	21.56
facesim	641128	554.326	545.10	568.44
ferret	106712	47.39	46.11	48.95
freqmine	99932	44.66	44.54	44.89
swaptions	29944	50.216	48.70	51.54
fluidanimate	323136	5966.668	4598.53	6864.11
vips	66968	47.724	47.61	47.89
x264	66976	29.416	29.21	29.60
canneal	232016	47.166	47.01	47.45
dedup	373812	1054.644	1026.09	1099.64
streamcluster	45844	58.51	58.49	58.55
raytrace	310600	512.842	512.42	513.72

Intel Thread Checker 3.1

Package	Contexts (avg)	Contexts (min)	Contexts (max)	Memory (avg)	Memory (min)
blackscholes	0	0	0	29426.4	25728
bodytrack	13	13	13	407736	311468
facesim	0	0	0	7582441.6	6874548
ferret	0	0	0	277602.4	263376
freqmine	1063.2	1060	1068	1718838.4	1716516
swaptions	0	0	0	25186.4	22152
fluidanimate	0	0	0	928003.2	926728
vips	0	0	0	698168.8	689360
x264	1	1	1	74528.8	66528
canneal	4	4	4	1499354.4	1490016
dedup	0	0	0	3198069.6	3168068
streamcluster	2	2	2	89924	83524
raytrace	0	0	0	4962412.8	4946100

Intel Thread Checker 3.1

Package	Memory (max)	Seconds (avg)	Seconds (min)	Seconds (max)
blackscholes	43536	19.846	15.78	21.49
bodytrack	469556	91.962	81.35	102.29
facesim	7772160	1346.868	1340.84	1358.71
ferret	325492	180.758	174.80	193.58
freqmine	1725064	278.68	247.38	312.29
swaptions	31232	518.918	499.12	527.21
fluidanimate	930788	245.598	241.05	256.77
vips	731648	373.922	357.34	386.38
x264	100836	0.788	0.66	1.10
canneal	1527076	102.222	96.14	105.13
dedup	3240488	296.608	288.17	302.64
streamcluster	96164	527.958	435.51	559.14
raytrace	5021720	1912.01	1900.89	1927.10

Helgrind 3.3.1

Package	**Contexts (avg)**	**Contexts (min)**	**Contexts (max)**	**Memory (avg)**	**Memory (min)**
blackscholes	2	2	2	55544	54512
bodytrack	223.6	222	225	71381.6	70468
facesim	112.6	109	121	787140.8	785076
ferret	111	111	111	134506.4	133348
freqmine	225.6	217	236	170896.8	170104
swaptions	0	0	0	56060.8	55748
fluidanimate	58	58	58	139136	138728
vips	69.4	65	75	87351.2	84940
x264	486.4	440	527	91676	90648
canneal	2	2	2	213903.2	213464
dedup	3	3	3	467438.4	462180
streamcluster	70	70	70	62383.2	59680
raytrace	117	117	117	551179.2	549988

Helgrind 3.3.1

Package	**Memory (max)**	**Seconds (avg)**	**Seconds (min)**	**Seconds (max)**
blackscholes	56624	11.444	11.38	11.49
bodytrack	71884	31.184	30.77	32.39
facesim	788704	971.146	948.67	987.22
ferret	138116	59.472	58.59	60.59
freqmine	171796	56.378	55.35	57.09
swaptions	56468	88.446	88.30	88.72
fluidanimate	139868	425.576	425.09	426.33
vips	91024	65.604	65.42	65.89
x264	93208	54.062	53.99	54.17
canneal	214656	112.124	110.13	113.11
dedup	471656	70.086	69.12	71.27
streamcluster	71176	119.57	118.67	120.03
raytrace	554356	912.308	911.23	913.83