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DEPARTMENT OF INTELLIGENT SYSTEMS
ÚSTAV INTELIGENTNÍCH SYSTÉMŮ

ADVANCED STATIC ANALYSIS OF ATOMICITY IN CONCURRENT PROGRAMS THROUGH FACEBOOK INFER

POKROČILÁ STATICKÁ ANALÝZA ATOMIČNOSTI V PARALELNÍCH PROGRAMECH V PROSTŘEDÍ FACEBOOK INFER

MASTER'S THESIS
DIPLOMOVÁ PRÁCE

AUTHOR **Bc. DOMINIK HARMIM**
AUTOR PRÁCE

SUPERVISOR **prof. Ing. TOMÁŠ VOJNAR, Ph.D.**
VEDOUCÍ PRÁCE

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Master's Thesis Specification

24185

Student: **Harmim Dominik, Bc.**
Programme: Information Technology and Artificial Intelligence
Specialization: Software Verification and Testing
Title: **Advanced Static Analysis of Atomicity in Concurrent Programs through Facebook Infer**
Category: Software analysis and testing
Assignment:

1. Study limitations of the atomicity analyser Atomer developed in your bachelor thesis as well as the latest developments concerning the Facebook Infer framework.
2. Propose ways of significantly improving precision and/or scalability of the analysis even if for the price of the user providing more input and/or combining it with dynamic analysis.
3. Implement a new version of Atomer including the proposed improvements and supporting analysis of programs written in more programming languages than just C supported by the first version of Atomer.
4. Evaluate the new version of Atomer on suitable benchmarks, including at least real-life code in which some atomicity problems were previously detected.
5. Describe and discuss the achieved results and their further possible improvements.

Recommended literature:

1. Rival, X., Yi, K.: Introduction to Static Analysis: An Abstract Interpretation Perspective. MIT Press, 2020.
2. Blackshear, S., Gorogiannis, N., O'Hearn, P. W., Sergey, I.: RacerD: Compositional Static Race Detection. In: Proc. of OOPSLA'18, PACMPL 2(OOPSLA):144:1-144:28, 2018.
3. Gorogiannis, N., O'Hearn, P.W., Sergey, I.: A True Positives Theorem for a Static Race Detector. In: Proc. of POPL'19, PACMPL 3(POPL):57:1-57:29, 2019.
4. Dias, R.J., Ferreira, C., Fiedor, J., Lourenço, J.M., Smrčka, A., Sousa, D.G., Vojnar, T.: Verifying Concurrent Programs Using Contracts, In: Proc. of ICST'17, IEEE, 2017.
5. Harmim, D.: Static Analysis Using Facebook Infer to Find Atomicity Violations. Bachelor thesis, Brno University of Technology, 2019.
6. Marcin, V.: Static Analysis Using Facebook Infer Focused on Deadlock Detection. Bachelor thesis, Brno University of Technology, 2019.

Requirements for the semestral defence:

- Item 1 and at least some development falling under items 2 and 3 of the assignment.

Detailed formal requirements can be found at https://www.fit.vut.cz/study/theses/

Supervisor: **Vojnar Tomáš, prof. Ing., Ph.D.**
Head of Department: Hanáček Petr, doc. Dr. Ing.
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Abstract

Atomer is a *static analyser* based on the idea that if some *sequences of functions* of a *multi-threaded program* are executed *under locks* in some runs, likely, they are *always intended to execute atomically.* Atomer thus strives to look for such sequences and then detects for which of them the atomicity may be broken in some other program runs. The author of this master's thesis proposed and implemented the first version of Atomer as a plugin of the *Facebook Infer framework* in his bachelor's thesis. In the master's thesis, a new and *significantly improved* version of Atomer is proposed. The improvements aim at both increasing *scalability* as well as *precision.* Moreover, support for several initially not supported programming features has been added (including, e.g., the possibility of analysing *C++ and Java programs* or support for *re-entrant locks* or *lock guards*). Through a number of experiments (including experiments with *real-life code* and *real-life bugs*), it is shown that the new version of Atomer is indeed much *more general, scalable, and precise.*

Abstrakt

Nástroj *Atomer* je *statický analyzátor* založený na myšlence, že pokud jsou některé *sekvence funkcí vícevláknového programu* prováděny v některých bězích *pod zámky*, je pravděpodobně zamýšleno, že mají být *vždy provedeny atomicky.* Analyzátor Atomer se tudíž snaží takové sekvence hledat a poté zjišťovat, pro které z nich může být v některých jiných bězích programu porušena atomicita. Autor této diplomové práce ve své bakalářské práci navrhl a implementoval první verzi nástroje Atomer jako zásuvný modul *aplikačního rámce Facebook Infer.* V této diplomové práci je navržena nová a *výrazně vylepšená* verze analyzátoru Atomer. Cílem vylepšení je zvýšení jak *škálovatelnosti*, tak *přesnosti.* Kromě toho byla přidána podpora pro několik původně nepodporovaných programovacích vlastností (včetně např. možnosti analyzovat *programy napsané v jazycích C++ a Java* nebo podpory pro *reentrantní zámky* nebo *stráže zámků*, tzv. „lock guards"). Prostřednictvím řady experimentů (včetně experimentů s *reálnými programy* a *reálnými chybami*) se ukázalo, že nová verze nástroje Atomer je skutečně *mnohem obecnější, přesnější a lépe škáluje*

Keywords

Facebook Infer, static analysis, abstract interpretation, contracts for concurrency, atomicity violation, concurrent programs, programs analysis, atomicity, incremental analysis, modular analysis, compositional analysis, interprocedural analysis, scalability, Atomer, function calls sequence, multi-threaded programs

Klíčová slova

Facebook Infer, statická analýza, abstraktní interpretace, kontrakty pro paralelismus, porušení atomicity, paralelní programy, analýza programů, atomicita, inkrementální analýza, modulární analýza, kompoziční analýza, interprocedurální analýza, škálovatelnost, Atomer, sekvence volání funkcí, vícevláknové programy

Reference

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Rozšířený abstrakt

[[Tady nějak zkombinovat a zkrátit úvod + závěr a přeložit do češtiny.]]

Advanced Static Analysis of Atomicity in Concurrent Programs through Facebook Infer

Declaration

Hereby I declare that this master's thesis was prepared as an original author's work under the supervision of professor Tomáš Vojnar. All the relevant information sources used during this thesis's preparation are appropriately cited and included in the reference list.

....
Dominik Harmim
28th April 2021

Acknowledgements

I thank my colleagues from VeriFIT for their assistance. I would particularly like to thank my supervisor Tomáš Vojnar. I also wish to acknowledge Nikos Gorogiannis from the Infer team at Facebook for valuable discussions about the analyser's development. Lastly, I acknowledge the financial support received from the H2020 ECSEL projects AQUAS, Arrowhead Tools, and VALU3S.

Contents

Chapter 1

Introduction

Bugs have been present in computer programs ever since the inception of the programming discipline. Unfortunately, they are often hidden in unexpected places, and they can lead to unexpected behaviour, which may cause significant damage. Nowadays, developers have many possibilities of catching bugs in the early development process. *Dynamic analysers* or tools for *automated testing* are often used, and they are satisfactory in many cases. Nevertheless, they can still leave too many bugs undetected because they can analyse only *particular program flows* dependent on the input data. An alternative solution is *static analysis* (despite it, of course, suffers from some problems too — such as the possibility of reporting many *false alarms*, i.e., *spurious errors*). Quite some tools for static analysis were implemented, e.g., Coverity or CodeSonar. However, they are often proprietary and difficult to openly evaluate and extend.

Recently, Facebook introduced *Facebook Infer*: an *open-source* tool for creating *highly scalable*, *compositional*, *incremental*, and *interprocedural* static analysers. Facebook Infer has grown considerably, but it is still under active development by many teams across the globe. It is employed every day not only in Facebook itself but also in other companies, such as Spotify, Uber, Mozilla, or Amazon. Currently, Facebook Infer provides several analysers that check for various types of bugs, such as buffer overflows, data races and some forms of deadlocks and starvation, null-dereferencing, or memory leaks. However, most importantly, Facebook Infer is a *framework* for building new analysers quickly and easily. Unfortunately, the current version of Facebook Infer still lacks better support for *concurrency bugs*. While it provides a reasonably advanced data race analyser, it is limited to Java and C++ programs only and fails for C programs, which use a *lower-level lock manipulation*. Moreover, the only available checker of *atomicity of call sequences* is the first version of *Atomer* [15] proposed in the bachelor's thesis of the author.

At the same time, in *concurrent programs*, there are often *atomicity requirements* for the execution of specific sequences of instructions. Violating these requirements may cause many kinds of problems, such as unexpected behaviour, exceptions, segmentation faults, or other failures. *Atomicity violations* are usually not verified by compilers, unlike syntactic or some sorts of semantic rules. Moreover, atomicity requirements, in most cases, are not even documented at all. Therefore, in the end, programmers themselves must abide by these requirements and usually lack any tool support. Furthermore, in general, it is difficult to avoid errors in *atomicity-dependent programs*, especially in large projects, and even more laborious and time-consuming is finding and fixing them. The papers [11, 13, 32, 35] discuss the importance of *atomicity-related bugs*, and they also show some bugs in *real-world programs*. Unfortunately, tool support for automatically discovering such kinds of errors is currently minimal.

As already mentioned, within the author's bachelor's thesis [15], *Atomer*[1] was proposed — a *static analyser* for finding some forms of *atomicity violations* implemented as a Facebook Infer's module. In particular, the stress is put on the *atomic execution of sequences of function calls*, which is often required, e.g., when using specific library calls. For example, assume the function `replace` from Listing 1.1 that replaces item `a` in an array by item `b`. It contains atomicity violation — the index obtained may be outdated when `set` is executed (because, e.g., a *concurrent thread* can modify the array), i.e., `index_of` and `set` should be executed atomically. The analysis is based on the assumption that sequences of function calls executed *atomically once* should probably be executed *always atomically*. Hence, the checker naturally works with sequences. In fact, the idea of checking the atomicity of certain sequences of function calls is inspired by the works of *contracts for concurrency* [11, 32]. In the terminology of [11, 32], the atomicity of specific sequences of calls is the most straightforward (yet very useful in practice) kind of contracts for concurrency. However, while the idea of using sequences in the given context is indeed natural and rather exact, it quite severely limits the *scalability* of the analysis (indeed, even with a few functions, there can appear numerous different orders in which they can be called). Moreover, the implementation of the first version of Atomer targets mainly *C programs* using *PThread locks*. Consequently, there was no support for other languages and their locking mechanisms in the first version of Atomer.

```
void replace(int a, int b, Array &array)
{
    int i = array.index_of(a);
    if (i >= 0) array.set(i, b);
}
```

Listing 1.1: An example of *atomicity violation*

Within this thesis, Atomer has been *significantly improved and extended*. In particular, to improve scalability, working with *sequences* of function calls was *approximated* by working with *sets* of function calls. Furthermore, several new features were implemented: support for *C++ and Java*, including various advanced kinds of *locks* these languages offer (such as *re-entrant locks* or *lock guards*); or a more precise way of *distinguishing between different lock instances*. Moreover, the analysis has been *parameterised* by function names to concentrate on during the analysis and limits of the number of functions in *critical sections*. These parameters aim to reduce the number of false alarms. Their proposal is based on the author's analysis of false alarms produced by the first Atomer's version. Lastly, new experiments were performed to test capabilities of the new version of Atomer.

The development of the original Atomer started under the H2020 ECSEL projects AQUAS and Arrowhead Tools. The development of its new version is supported by the H2020 ECSEL project VALU3S. It has been discussed with the developers of Facebook Infer too.
Parts of the thesis concerning the preliminaries and the basic version of Atomer are partially

[1]The implementation of **Atomer** is available at GitHub as an *open-source* repository (in a branch `atomicity-sets`): `https://github.com/harmim/infer`.

taken from the thesis [15]. Moreover, some preliminary results were also published in the Excel@FIT paper [16] written by the author.

The rest of the thesis is organised as follows. Chapter 2 describes all the topics related to and essential to this thesis (including *static analysis*, *abstract interpretation*, *Facebook Infer*, and *contracts for concurrency*). The original version of *Atomer*, its limitations, and related work are described in Chapter 3. Subsequently, Chapter 4 presents all the proposed extensions and improvements. The implementation of these extensions is then covered in Chapter 5. The experimental evaluation of the new Atomer's features and other experiments performed within this thesis are discussed in Chapter 6 together with the future work. Finally, the thesis is concluded in Chapter 7. Besides, there are included the following appendices. The content of the attached memory media is listed in Appendix A. In the end, Appendix B serves as an installation and user manual.

Chapter 2

Preliminaries

This chapter explains the theoretical background that the thesis builds on. It also explains and describes the existing tools used in the thesis. Also, note that the author already partially published the contents of the sections in this chapter in his bachelor's thesis [15] and paper [16].

Multi-Threaded Programs [13, 25] *Multiple threads* of control are commonly used in the software development process because they help reduce *latency*, increase *throughput*, and better utilise *multiple processor computers*. Threads or *processes* are independent *sequences of instructions* that may be performed simultaneously. A process represents a single running program. It has its own address space and a unique identifier. One process can consist of multiple threads, i.e., a thread is a so-called *lightweight process*. All threads of a single process share the same address space (i.e., code and data). However, reasoning about the behaviour and correctness of a multi-threaded system is complex due to the need to consider all possible interleavings of the executions of different threads. An integral part of parallel programming is the *synchronisation* of individual threads. Usually, synchronisation mechanisms ensure the *mutual exclusion* of shared resources or synchronise actions that threads perform. Operating systems provide basic synchronisation primitives that can be often used in programming languages in a higher-level manner. Such fundamental mechanisms are *semaphores* (binary semaphores are called *mutexes*), *barriers*, *read-copy-update* techniques, or *monitors*. In this work (and also in practice), for simplicity, it will be used a notion of *locks* for semaphores, mutexes, monitors, etc. as a mechanism that mutually excludes access to a *critical section*.

The use of locks and access to shared data in parallel programs involves the risk of errors not known in sequential programming. Much previous work on checking *thread interference* has focused on *data races* (or, in general, on *race conditions*). A data race occurs when two threads simultaneously access the same data variable, and, at least, one of the access is a write. In practice, data races are commonly avoided by guarding each data structure with a lock. Unfortunately, the absence of data races is insufficient to guarantee the absence of errors due to unexpected interference between threads. [13, 35]

Atomicity Violations [13] Another possible source of errors due to unexpected interference between threads are *atomicity violations*. A method (or, in general, a code block) is *atomic* iff, for every (arbitrarily interleaved) program execution, there is an equivalent execution with the same overall behaviour where the atomic method is executed serially. In other words, the method's execution is not interleaved with instructions of other threads. Also, atomicity provides a powerful, indeed maximal, guarantee of non-interference among threads. In short, atomicity is a generally applicable and fundamental *correctness property* of multi-threaded code. Nevertheless, commonly used *testing approaches* are deficient in verifying atomicity. While testing may reveal a concrete interleaving in which an atomicity violation causes erroneous behaviour, the exponentially large amount of possible interleavings does fundamentally impossible to get suitable *test coverage*. Research results [11, 13, 32, 35] have shown that defects related to atomicity are common, even in well-tested libraries.

It has to be described which blocks of code should be executed atomically to detect atomicity violations. Algorithms for the detection of this error often deal only with the atomicity of operations on variables. Examples of a complex description of *atomicity requirements* are *contracts* that require atomicity of the execution of certain methods/functions (defined in Section 2.4). [25]

Further, Section 2.1 outlines the fundamental notions and approaches in *program analysis*. In Section 2.2, there is an explanation of *static analysis* by *abstract interpretation*, which is used in *Facebook Infer*, i.e., the key framework used in this thesis. Facebook Infer, its principles, and features are covered in Section 2.3. The concept of *contracts for concurrency* is discussed and defined in Section 2.4.

2.1 Concepts in Program Analysis

This section provides a fundamental intuition about the main principles of *program analysis*. It discusses several standard techniques for reasoning about programs. The section is based on a few first chapters from the book [30] and the overview paper [19].

2.1.1 What to Analyse

The first question is *what programs to analyse*. An obvious characterisation of programs to analyse is the programming language in which the programs are written. Moreover, certain specific families of input programs may be distinguished. However, besides the language and the family of programs to consider, the way input programs are processed can also differ and affects how the analysis works. The most straightforward way to handle input programs is working directly with a *source code* just like a compiler would. Nevertheless, different representations of programs are used in program analysis likewise. The two below types of techniques are customarily distinguished.

Program-Level Analysis The first possibility is to run the analysis on a *source code* of input programs (e.g., programs written in conventional programming languages like C or Java; or hardware described in VHDL, Verilog, etc.) or on *executable binaries*. This technique typically involves some *front-end* comparable to a compiler's one that constructs the *syntax trees* of input programs.

Model-Level Analysis An alternative option is to consider a different input language that aims at modelling the semantics of programs. Then, the analysis takes as an input a description that *models* the program to analyse. Such models either need to be created manually, or some specialised tools are used. These models may hide implementation difficulties or inaccuracies. Examples of the models are automata of any kind, UML diagrams, Petri nets, Markov chains, or specialised modelling languages like, e.g., Promela.

2.1.2 Static vs Dynamic Approaches

Another critical question in program analysis is *when* the analysis is made. In particular, whether it operates *during* or *before* the program's execution.

One solution is to analyse programs at *run-time*, i.e., *during* the program's execution. This approach is called *dynamic analysis*. It takes place while the program computes, often, over a number of executions.

A second approach is to analyse the program *before* execution, which is called *static analysis*. It is done independently from any execution.

Static and dynamic approaches are significantly distinct in many ways. They come with different benefits and weaknesses. While dynamic approaches are often simpler to design and implement, they often have problems with performance at run-time. They do not force developers to fix bugs before program execution. Moreover, some properties cannot be checked (or, at least, it is very challenging) dynamically. For instance, dynamically detecting whether an execution does not *terminate* would require an infinite program run. Particular static and dynamic techniques are further discussed in Section 2.1.5.

2.1.3 Automation and Scalability

Automation is a further relevant aspect in program analysis. It would be ideal if program analysis methods were *fully automated* (i.e., no human help is needed). Unfortunately, this is not always possible due to the consequences of *Rice's theorem* [29]. Thus, sometimes, it is needed to give up on automation and let program analyses ask some *user input* (i.e., some human help is required). In that case, the user is asked to give some information to the analysis, e.g., *invariants*[1]. That is, the analysis is partially manual since users need to compute parts of the results themselves. However, having to provide this information may be unwieldy because input programs can be huge or complex.

Scalability is another essential characteristic of program analysis algorithms. Even if a program analysis is fully automatic, it is not guaranteed that it will generate a result within a reasonable time, depending on the complexity of the algorithms. A program analysis tool may not *scale* to extensive programs due to time costs or, e.g., memory constraints.

2.1.4 Soundness and Completeness

In order to preserve automation and/or scalability, the conditions about program analysis may be relaxed. Namely, the analysis can be proposed to return *inaccurate* results (altern-

[1]An **invariant** is a *logical property* that can be proved to be inductive for a given program.

atively, it can return a non-conclusive "do not know" answer). For this purpose, two dual properties (forms of *approximations* or *inaccuracies*) are used. To express these notions, let $\mathcal{L}$ be a *Turing-complete language*, φ be a *non-trivial semantic property* of interest of programs of $\mathcal{L}$, and T be an analysis tool that decides whether φ holds in a given program. Ideally, if T were absolutely precise, it would be such that:

$$\text{for every program } P \in \mathcal{L} : T(P) = \mathit{true} \iff P \models \varphi$$

The above, of course, can be decomposed into a pair of implications:

$$\begin{cases} \forall P \in \mathcal{L} : T(P) = \mathit{true} \implies P \models \varphi \\ \forall P \in \mathcal{L} : T(P) = \mathit{true} \impliedby P \models \varphi \end{cases}$$

Soundness A *sound* program analysis satisfies the first implication.

Definition 2.1.1 (**Soundness** [30])**.** The program analyser T is sound w.r.t. property φ whenever, for any program $P \in \mathcal{L}$, $T(P) = \mathit{true} \implies P \models \varphi$.

When a sound analysis terminates and claims that the analysed program has property φ, it guarantees that the program indeed satisfies φ, i.e., no errors are missed. In other words, a sound analysis will refuse all programs that do not satisfy φ. In terms of errors in *binary classification*, there are not *false-negative* errors[2] when using sound analysis. On the other hand, there is a chance of *false-positive* errors[3].

Completeness A *complete* program analysis satisfies the second implication.

Definition 2.1.2 (**Completeness** [30])**.** The program analyser T is complete w.r.t. property φ whenever, for any program $P \in \mathcal{L}$, $P \models \varphi \implies T(P) = \mathit{true}$.

A complete program analysis will accept every program that satisfies property φ. In other words, when a complete analysis refuses an analysed program, it is guaranteed that the program indeed fails to satisfy φ, i.e., there are no false-positive errors. However, there can be false-negative errors.

Due to the *computability* barrier [29], it is not possible to design a general analysis to determine which programs satisfy any non-trivial property for a Turing-complete language that is sound, complete, fully automated, and scalable at the same time. Some of these have to be sacrificed, or the analysis has to be proposed to operate only on a specific set of input programs.

2.1.5 Program Analysis Techniques

This section describes several standard *program analysis techniques*. Although this thesis focuses on *static analysis* (or maybe rather *bug finding* based on static analysis), it is essential to see the differences between other analysis techniques. Hence, a brief overview is presented in the below sections. Finally, Table 2.1 compares the techniques (how they are usually used) based on the earlier criteria.

[2] A **false-negative** error is a real error that is undetected by an analysis tool.

[3] A **false-positive** error (also called a **false alarm**) is a *spurious error*, i.e., it is detected by an analysis tool, but the error does not exist in the real program.

Testing and Dynamic Analysis The *testing* approach checks a finite set of finite program executions. In the development process, several levels of testing are used at various stages of the software/hardware life-cycle, e.g., *unit testing* or *integration testing.* In general, it is challenging to achieve suitable *test coverage* because of infinite program paths when using, e.g., *random testing.* However, several more advanced techniques that improve the coverage have been introduced. These techniques are often combined with other verification approaches. For instance, *concolic testing* combines testing with *symbolic execution*, *dynamic analysis* that observes behaviour in a testing run (such behaviour can be *extrapolated* to behaviour not seen in the given testing run), or *search-based techniques* that can generate test data or parameters. Moreover, to test programs with non-deterministic semantics (i.e., concurrent programs), techniques like *noise injection* are applied. In particular, many advanced dynamic analysers have been proposed to detect data races or deadlocks, including, e.g., Eraser or FastTrack.

Testing has the following features. In general, it is simple to automate. It is unsound in almost all cases (besides programs that have a finite number of finite paths). Since failed testing runs provide incorrect concrete executions, the testing is complete.

Other analysis techniques mentioned below use a static approach. The significant difference between static and dynamic approaches is the following. It is well-known that testing may expose errors, but it cannot prove their absence. It was also famously stated by Edsger W. Dijkstra: "*Program testing can be used to show the presence of bugs, but never to show their absence!*". However, static approaches may be able to prove their absence — with some *approximation* — they can check *all possible executions* of a program and provide guarantees about its properties. Another static approaches' benefit is that the analysis can be performed during the development process, so the program does not have to be executable yet, and it already can be analysed. The biggest drawback of static approaches, in general, is that they can produce a lot of false alarms (though this is not the case, e.g., for *theorem proving*), which is often resolved by accepting unsoundness. Another crucial issue of static approaches (this is, however, also an issue of dynamic analyses) is ensuring sufficient scalability — in fact, typically, the more precise the analysis, the less scalable it becomes.

Deductive Verification The *semi-automated* approach *deductive verification* (or *theorem proving*) uses *inference systems* for inferring theorems about the analysed system from the facts known about the system and from general theorems of different logical theories. The approach falls under *machine-assisted* techniques, which means that users may be required to provide extra information to the analysis (usually *loop invariants*, procedure *preconditions/postconditions*, assertions, or some other invariants). This can be demanding and require some level of expertise. However, a substantial part of the verification can usually still be carried out in a fully automated fashion. This approach is very general, but there is a problem with generating diagnostic information for incorrect systems. There exist a number of *interactive theorem proving* tools like Coq, Isabelle/HOL, PVS, ACL2, etc. The user usually guides the inference process in these tools.

These techniques also involve *automated decision procedures* (or *satisfiability solvers*) for various logical theories. Such solvers are often used as back-end components for higher-level verification methods, such as *symbolic execution* or *predication abstraction* in model checking. Commonly used solvers are *SAT-solvers* (e.g., CaDiCaL and Glucose) and *SMT-*

solvers (e.g., Z3 and CVC4). Various tools allow the user to provide some *logical annotations* in a code and then automatically attempt to prove specific properties using decision procedures. Examples of such tools are VCC and Dafny.

Theorem proving techniques have the following properties. They are not fully automatic, i.e., *high user expertise* is often needed. They are sound w.r.t. the model of the program used in the proof, and they are usually complete up to the capabilities of the proof assistant.

Model Checking Another technique called (*finite-state*) *model checking* aims at finite systems. It automatically verifies whether a system or its model satisfies a particular property based on an *algorithmic exploration of the system's state space.* Unfortunately, the biggest issue here is the *state space explosion* problem [33]. However, in practice, model checking tools use effective data structures (such as *binary decision diagrams* or *hierarchical storage of states*) to describe program behaviours and avoid enumerating all executions thanks to approaches that reduce the search space. In addition, other techniques are used to cope with this problem. For instance, various abstractions are used (e.g., *predicate abstraction*) or *bounded model checking*[4]. Properties are usually defined using *temporal logics*, such as LTL, CTL, or PCTL. Finite model checking has the following characteristics. It is automatic (up to the need of modelling the system or its environment). It is sound and usually complete, w.r.t. the model. Other advantages are that it is pretty general and provides diagnostic data for incorrect systems.

Model checking is typically done at the *model level*, i.e., a model of the program needs to be built, either manually or automatically. In practice, model checking tools usually implement a front-end for that purpose. The problem is that the model generally cannot precisely capture the input program's behaviours since programs are usually infinite systems. Thus, the checking of the synthesised model may be either incomplete or unsound, w.r.t. the input program. Some model checking techniques can automatically *refine* the model when they fail to prove a property due to a *spurious counterexample*, although a *termination* must be ensured. In practice, model checking tools are often sound and incomplete w.r.t. the input program.

Model checking has found many successful applications, including hardware verification, verification of *concurrent* and *distributed* systems, *probabilistic* systems, *biological* systems, etc. Examples of hardware model checkers are RuleBase, Incisive Verifier, or NuSMV. Model checkers for concurrent and distributed systems include Spin or DIVINE. Tools, such as CPAchecker or BLAST, are model checkers that use predicate abstraction. PRISM and Storm are state-of-the-art tools for model checking probabilistic systems. Finally, Uppaal is a model checker that verifies temporal logic formulas on *timed automata*.

Static Analysis ...

[4]**Bounded model checking** explores models up to a *fixed depth* only. This technique is often referred to as *bug finding* because it sacrifices completeness and often also soundness. Examples of tools that implement such technique are CMBC, LLBMC, or JBMC.

Table 2.1: A summary of *program analysis techniques* [30]

Technique	Automatic	Sound	Complete	Object	Approach
Testing	Yes	No	Yes	Program	Dynamic
Deductive verification	No	Yes	Yes/No	Model/Program	Static
Model checking	Yes	Yes	Yes/No	Model/Program	Static
Conservative static analysis	Yes	Yes	No	Program	Static
Bug finding	Yes	No	No	Program	Static

2.2 Abstract Interpretation

[[Přidat Galoisovo spojení. Definovat widening, narrowing.]]

According to [26], *static analysis* of programs is reasoning about the behaviour of computer programs without actually executing them. It has been used since the 1970s in optimising compilers for generating efficient code. More recently, it has proven valuable also for automatic error detection, verification of correctness of programs, and it is used in other tools that can help programmers. Intuitively, a static program analyser is a program that reasons about the behaviour of other programs, in other words, a static program analyser is a program that reasons about another programs by looking for some *syntactic patterns* in the code and/or by assigning the program statements some *abstract semantics* and then deriving a characterisation of the behaviour in terms of the abstract semantics. Nowadays, static analysis is one of the fundamental concepts of *formal verification*. It aims to automatically answer questions about a given program, such as, e.g., [26]:

- **Are certain operations executed *atomically*?**
- Does the program terminate on every input?
- Can the program *deadlock*?
- Does there exist an input that leads to a *null-pointer dereference*, a *division-by-zero*, or an *arithmetic overflow*?
- Are all variables initialised before they are used?
- Are arrays always accessed within their bound?
- Does the program contain *dead code*?
- Are all resources correctly released after their last use?

Various forms of static analysis of programs have been invented, for instance [15]: bug pattern searching, data-flow analysis, constraint-based analysis, type analysis, or symbolic execution. One of the most widely used approaches — *abstract interpretation* — is detailed in Section 2.2.

There exist numerous tools for static analysis (often proprietary and difficult to openly evaluate or extend), e.g.: Coverity, Klockwork, CodeSonar, Frama-C, PHPStan, or *Facebook Infer* (described in Section 2.3).

2.2.1 Abstract Interpretation

[[Pro definici summary citovat Hoareho a použít jeho trojice.]]

This section explains and defines the basics of *abstract interpretation*. The description is based on [8, 9, 7, 6, 20, 15, 10, 27, 26]. In these works, there can be found more detailed and more formal explanation.

Abstract interpretation was introduced and formalised by a French computer scientist Patrick Cousot and his wife Radhia Cousot in the year 1977 at POPL (symposium on Principles of Programming Languages) [9]. It is a generic *framework* for static analyses. It allows one to create particular analyses by providing specific components (described later) to the framework. The obtained analysis is guaranteed to be *sound* if certain properties of the components are met. [20, 15]

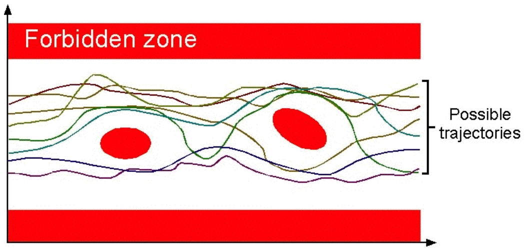
In general, in the set theory, which is independent of the application setting, abstract interpretation is considered a theory for *approximating* sets and set operations. A more restricted formulation of abstract interpretation is to interpret it as a theory of approximation of the behaviour of the *formal semantics* of programs. Those behaviours may be characterised by *fixpoints* (defined below), which is why a primary part of the theory provides efficient techniques for *fixpoint approximation* [27]. So, for a standard semantics, abstract interpretation is used to derive the approximate abstract semantics over an *abstract domain* (defined below). The abstract semantics obtained as a result of program analysis can then be used for verification, optimisation, code generation or transformation, etc. [8]

Patrick Cousot intuitively and informally illustrates abstract interpretation in [7] as follows. Figure 2.1a shows the *concrete semantics* of a program by a set of curves, which represents the set of all possible executions of the program in all possible execution environments. Each curve shows the evolution of the vector $x(t)$ of input values, state, and output values of the program as a function of the time t. *Forbidden zones* on this figure represent a set of erroneous states of the program execution. Proving that the intersection of the concrete semantics of the program with the forbidden zone is empty may be *undecidable* because the program concrete semantics is, in general, *not computable*. As demonstrates Figure 2.1b, abstract interpretation deals with an *abstract semantics*, i.e., the *superset* of the concrete program semantics. The abstract semantics includes all possible executions. That implies that if the abstract semantics is safe (i.e. it does not intersect the forbidden zone), the concrete semantics is safe as well. However, the *over-approximation* of the possible program executions causes that inexisting program executions are considered, which may lead to *false alarms*. It is the case when the abstract semantics intersects the forbidden zone, whereas the concrete semantics does not intersect it.

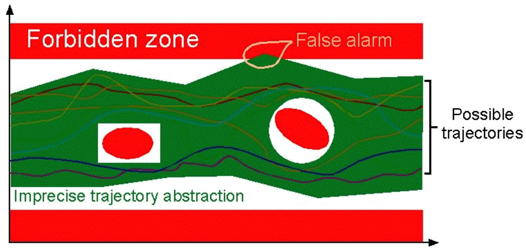
Components of Abstract Interpretation

In accordance with [20, 15], the basic components of abstract interpretation are as follows:

- **An Abstract Domain** [6]:



(a) *Concrete semantics* of programs with *forbidden zones*



(b) *Abstract semantics* of programs with imprecise trajectory abstraction

Figure 2.1: Abstract interpretation demonstration [7]. Horizontal axes: time t. Vertical axes: vector $x(t)$ of input, state, and output values of the considered program

  - An abstraction of the possible concrete program states (or their parts) in the form of *abstract properties*[5] and *abstract operations*[6] [8].
  - Sets of program states at certain locations are represented using *abstract states*.

- **Abstract Transformers**:
  - There is a *transform function* for each program operation (instruction) that represents the impact of the operation executed on an abstract state.

- **The Join Operator** $\sqcup$:
  - Joins abstract states from individual program branches into a single one.

- **The Widening Operator** $\triangledown$ [27, 10, 20]:
  - Enforces termination of the abstract interpretation.
  - It is used to approximate the *least fixed points* of program semantics (it is performed on a sequence of abstract states at a certain location).
  - Usually, the later in the analysis this operator is used, the more accurate the result is (but the analysis takes more time).

- **The Narrowing Operator** $\triangle$ [27, 10, 20]:
  - Using this operator, the approximation obtained by widening can be refined, i.e., it may be used to refine the result of widening.
  - This operator is used when a *fixpoint* is approximated using widening.

Fixpoints and Fixpoint Approximation

A **fixpoint** of a function $f : A \to A$ is an element $a \in A$ if and only if $\boldsymbol{f(a) = a}$ [15].

[5] **Abstract properties** approximating *concrete properties behaviours*.

[6] **Abstract operations** include abstractions of the *concrete approximation*, an approximation of the *concrete fixpoint transform function*, etc.

Computation of the *most precise abstract fixpoint* is not generally guaranteed to terminate, in particular, when a given program contains a loop or recursion. The solution is to approximate the fixpoint using *widening* (over-approximation of a fixpoint) and *narrowing* (improves the approximation of the fixpoint) [20, 15]. Most program properties can be represented as fixpoints. This reduces program analysis to the fixpoint approximation [6]. Further information about fixpoint approximation can be found, e.g., in [27, 10].

Formal Definition of Abstract Interpretation

According to [9, 20], **abstract interpretation** $\boldsymbol{I}$ of a program P with the instruction set S is a tuple

$$\boldsymbol{I} = (\boldsymbol{Q}, \sqcup, \sqsubseteq, \top, \bot, \boldsymbol{\tau})$$

where

- $\boldsymbol{Q}$ is the *abstract domain* (domain of *abstract states*),
- $\sqcup : Q \times Q \to Q$ is the *join operator* for accumulation of abstract states,
- $(\sqsubseteq) \subseteq Q \times Q$ is an *ordering* defined as $x \sqsubseteq y \Leftrightarrow x \sqcup y = y$ in $(Q, \sqcup, \top)$,
- $\top \in Q$ is the *supremum* of Q,
- $\bot \in Q$ is the *infimum* of Q,
- $\boldsymbol{\tau} : S \times Q \to Q$ defines *abstract transformers* for specific instructions,
- $(Q, \sqcup, \top)$ is a *complete semilattice* [15, 20].

Using so-called *Galois connections* [27, 10, 20, 6], one can guarantee the *soundness* of abstract interpretation.

2.3 Facebook Infer — Static Analysis Framework

[[Citovat Vladovu bakalářku. Zmínit se o C#.]]

This section describes the principles and features of *Facebook Infer*. The description is based on information provided at the Facebook Infer website[7] and in [2, 15]. Parts of this section are taken from the paper [15, 16].

Facebook Infer is an open-source[8] static analysis *framework*, which is able to discover various kinds of software bugs of a given program, with the stress put on *scalability*. The basic usage of Facebook Infer is illustrated in Figure 2.2. A more detailed explanation of its architecture is shown below. Facebook Infer is implemented in *OCaml*[9] – *functional* programming language, also supporting *imperative* and *object-oriented* paradigms. Further

[7] **Facebook Infer** website – https://fbinfer.com.
[8] **Open-source repository of Facebook Infer** on GitHub – https://github.com/facebook/infer.
[9] **OCaml** website – https://ocaml.org.

details about OCaml can be found in [24] or in official documentation[10], tutorials[11]. Facebook Infer was originally a standalone tool focused on *sound verification* of the absence of *memory safety violations*, which was first published in the well-known paper [5].

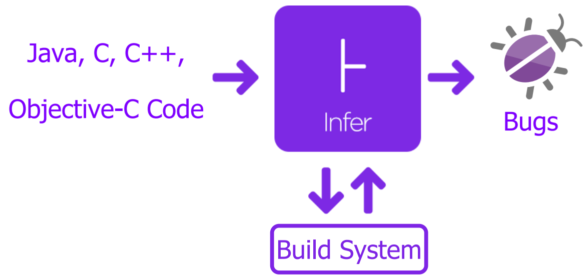


Figure 2.2: Static analysis in Facebook Infer (inspired by https://adtmag.com/articles/2015/06/12/facebook-infer.aspx)

Facebook Infer is able to analyse programs written in several languages. In particular, it supports languages C, C++, Java, and Objective-C. Moreover, it is possible to extend Facebook Infer's *frontend* for supporting other languages. Currently, Facebook Infer contains many analyses focusing on various kinds of bugs, e.g., *Inferbo* (buffer overruns) [36]; *RacerD* (data races) [3, 4, 14]; and other analyses that check for buffer overflows, some forms of deadlocks and starvation, null-dereferencing, memory leaks, resource leaks, etc.

2.3.1 Abstract Interpretation in Facebook Infer

Facebook Infer is a general framework for static analysis of programs, it is based on *abstract interpretation*. Despite the original approach taken from [5], Facebook Infer aims to find bugs rather than formal verification. It can be used to quickly develop new sorts of *compositional* and *incremental* analysers (*intraprocedural* or *interprocedural* [27]) based on the concept of function *summaries*. In general, a *summary* is a representation of *preconditions* and *postconditions* of a function. However, in practice, a summary is a custom data structure that may be used for storing any information resulting from the analysis of particular functions. Facebook Infer generally does not compute the summaries in the course of the analysis along the *Control Flow Graph* (**CFG**)[12] as it is done in classical analyses based on the concepts from [28, 31]. Instead, Facebook Infer performs the analysis of a program *function-by-function along the call tree*, starting from its leafs (demonstrated later). Therefore, a function is analysed and a summary is computed without knowledge of the call context. Then, the summary of a function is used at all of its call sites. Since summaries do not differ for different contexts, the analysis becomes more scalable, but it can lead to a loss of accuracy. In order to create a new intraprocedural analyser in Facebook Infer, it is needed to define the following (listed items are described in more detail in Section 2.2):

1. The *abstract domain* Q, i.e., a type of an *abstract state*.
2. Operator $\sqsubseteq$, i.e., an *ordering* of abstract states.

[10] **OCaml documentation** – http://caml.inria.fr/pub/docs/manual-ocaml.

[11] **OCaml tutorials** – https://ocaml.org/learn/tutorials.

[12] **A control flow graph (CFG)** is a directed graph in which the nodes represent basic blocks and the edges represent control flow paths [1]. **[[Křena/Vojnar - strana 6, poznámka pod čarou.]]**

3. The *join* operator $\sqcup$, i.e., the way of joining two abstract states.
4. The *widening* operator $\triangledown$, i.e., the way how to enforce termination of the abstract interpretation of an iteration.
5. *Transfer functions* τ, i.e., a transformer that takes an abstract state as an input and produces an abstract state as an output.

Further, in order to create an interprocedural analyser, it is required to additionally define:

1. The type of function summaries.
2. The logic for using summaries in transfer functions, and the logic for transforming an intraprocedural abstract state to a summary.

An important feature of Facebook Infer improving its scalability is *incrementality* of the analysis, it allows one to analyse separate code changes only, instead of analysing the whole codebase. It is more suitable for extensive and variable projects, where ordinary analysis is not feasible. The incrementality is based on *re-using summaries* of functions for which there is no change in them neither in the functions transitively invoked from them.

The Architecture of the Abstract Interpretation Framework in Facebook Infer

The architecture of the abstract interpretation framework of Facebook Infer (**Infer.AI**) may be split into three major parts, as demonstrated in Figure 2.3: a *frontend*, an *analysis scheduler* (and a *results database*), and a set of *analyser plugins*.

The frontend compiles input programs into the *Smallfoot Intermediate Language* (SIL) and represents them as a CFG. There is a separate CFG representation for each analysed function. Nodes of this CFG are formed as instructions of SIL. The SIL language consists of the following underlying instructions:

- `LOAD`: reading into a temporary variable.
- `STORE`: writing to a program variable, a field of a structure, or an array.
- `PRUNE e` (often called `ASSUME`): evaluation of a condition `e`.
- `CALL`: a function call.

The frontend allows one to propose *language-independent* analyses (to a certain extent) because it supports input programs to be written in multiple languages.

The next part of the architecture is the scheduler, which defines the order of the analysis of single functions according to the appropriate *call graph*[13]. The scheduler also checks if it is possible to analyse some functions simultaneously, which allows Facebook Infer to run the analysis in parallel.

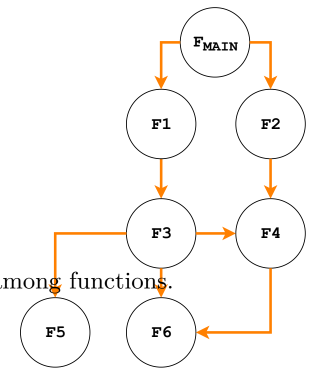


Figure 2.4: A call graph for an illustration of Facebook Infer's analysis process [2, 15]

[13] **A call graph** is a *directed graph* describing call dependencies among functions.

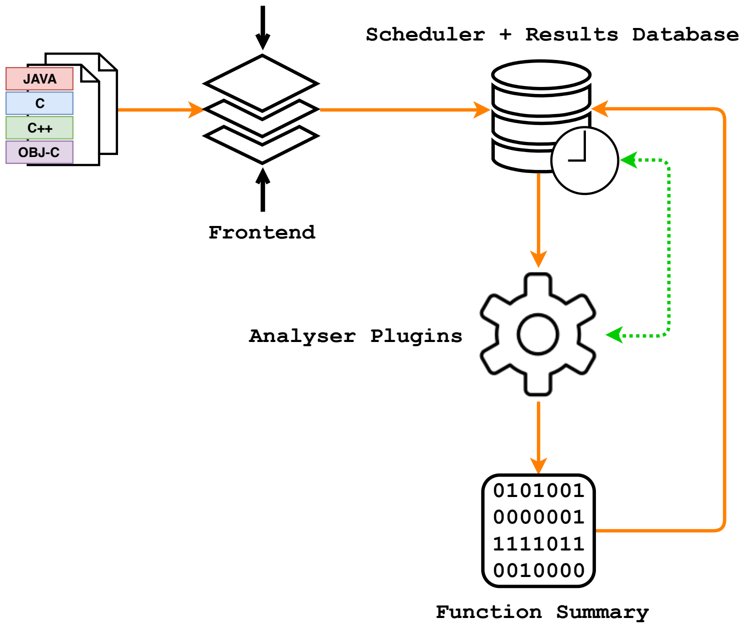


Figure 2.3: The architecture of Facebook Infer's abstract interpretation framework [2, 15]

Example 2.3.1. For demonstrating the order of the analysis in Facebook Infer and its incrementality, assume a call graph in Figure 2.4. At first, leaf functions F5 and F6 are analysed. Further, the analysis goes on towards the root of the call graph – $\mathtt{F}_{\mathtt{MAIN}}$, while taking into consideration the dependencies denoted by the edges. This order ensures that a summary is available once a nested function call is abstractly interpreted within the analysis. When there is a subsequent code change, only directly changed functions and all the functions up the call path are re-analysed. For instance, if there is a change of source code of function F4, Facebook Infer triggers re-analysis of functions F4, F2, and $\mathtt{F}_{\mathtt{MAIN}}$ only.

The last part of the architecture consists of a set of analyser plugins. Each plugin performs some analysis by interpreting of SIL instructions. The result of the analysis of each function (function summary) is stored to the results database. The interpretation of SIL instructions (*commands*) is done using an *abstract interpreter* (also called a *control interpreter*) and *transfer functions* (also called a *command interpreter*). The transfer functions take an actual *abstract state* of an analysed function as an input, and by applying the interpreting command, produce a new abstract state. The abstract interpreter interprets the command in an *abstract domain* according to the CFG. This workflow is shown in a simplified form in Figure 2.5.

2.4 Contracts for Concurrency

[[Přidat něco z prací Moniky Mužíkovské.]]

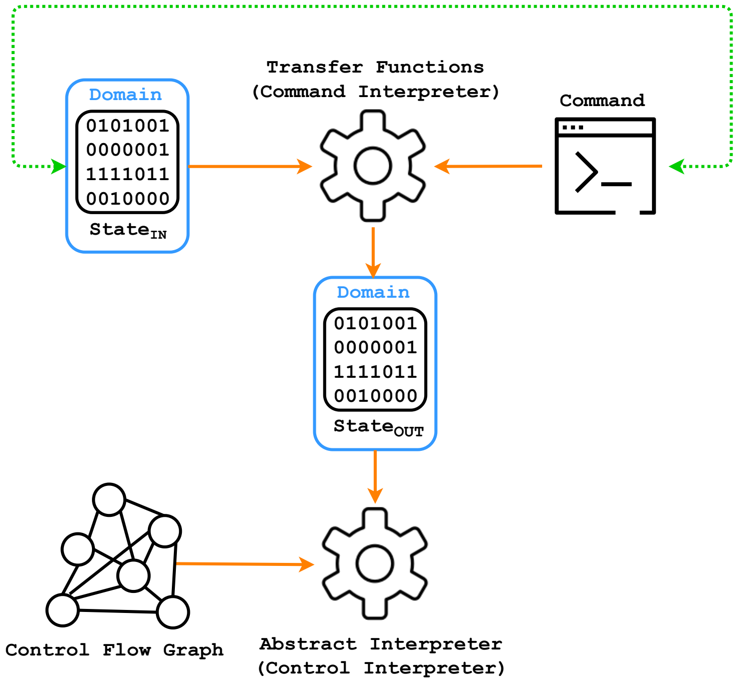


Figure 2.5: Facebook Infer's abstract interpretation process [2, 15]

This section introduces the concept of *contracts for concurrency* described in [32, 11]. Parts of this section are taken from the thesis [15]. Listings in this section are pieces of programs written in ANSI C.

Respecting the *protocol* of a software module — defines which *sequences of functions* are legal to invoke — is one of the requirements for the correct behaviour of the module. For example, a module that deals with a file system typically requires that a programmer using this module should call function `open` at first, followed by an optional number of functions `read` and `write`, and at last, call function `close`. A program utilising such a module that does not follow this protocol is erroneous. The methodology of *design by contract* (described in [23]) requires programs to meet such well-defined behaviours. [32]

In *concurrent programs*, contracts for concurrency allow one — in the simplest case — to specify *sequences of functions* that are needed to be *executed atomically* in order to avoid *atomicity violations*. In general, contracts for concurrency specify sets of sequences of calls that are called *spoilers* and sets of sequences of calls that are called *targets*, and it is then required that no target overlaps fully with any spoiler. Such contracts may be manually specified by a developer or they may be automatically generated by a program (analyser). These contracts can be used to verify the correctness of programs as well as they can serve as helpful documentation.

Section 2.4.1 defines the notion of *basic contracts for concurrency*. Further, Section 2.4.2 defines contracts extended to consider the *data flow* between functions (where a sequence of function calls must be atomic only if they handle the same data). The above mentioned more general contracts for concurrency with spoilers and targets, which essentially extend the basic contracts with some *contextual information*, are not presented here in detail (they are explained in the paper [11]). The reason is that the proposed analyser — *Atomer* — so far concentrates on the basic contracts.

2.4.1 Basic Contracts for Concurrency

In [11, 32], a *basic contract* is formally defined as follows. Let $\Sigma_{\mathbb{M}}$ be a set of all function names of a software module. A contract is a set $\mathbb{R}$ of *clauses* where each clause $\varrho \in \mathbb{R}$ is a *star-free regular expression*[14] over $\Sigma_{\mathbb{M}}$. A *contract violation* occurs if any of the sequences expressed by the contract clauses are interleaved with the execution of functions from $\Sigma_{\mathbb{M}}$, in other words, each sequence specified by any clause ϱ must be executed atomically, otherwise, there is a violation of the contract. The number of sequences defined by a contract is finite since the contract is the union of *star-free languages*.

Example 2.4.1. Consider the following example from [11, 32]. Assume that there is a module implementing a resizable array implementing the following interface functions:

f_1: `void add(char *array, char element)`

f_2: `bool contains(char *array, char element)`

f_3: `int index_of(char *array, char element)`

f_4: `char get(char *array, int index)`

f_5: `void set(char *array, int index, char element)`

f_6: `void remove(char *array, int index)`

f_7: `int size(char *array)`

The module's contract contains the following clauses:

(ϱ_1) `contains index_of`

> The execution of `contains` followed by the execution of `index_of` should be atomic. Otherwise, the program may fail to get the index, because after verification of the presence of an element in an array, it can be removed by some *concurrently running process*.

(ϱ_2) `index_of (get | set | remove)`

> The execution of `index_of` followed by the execution of `get`, `set`, or `remove` should be atomic. Otherwise, the received index may be outdated when it is applied to the address of an element, because a concurrent modification of an array may shift the position of the element.

(ϱ_3) `size (get | set | remove)`

[14] **Star-free regular expressions** are regular expressions using only the *concatenation operators* and the *alternative operators* (|), without the *Kleene star operator* (∗).

The execution of `size` followed by the execution of `get`, `set`, or `remove` should be atomic. Otherwise, the size of an array may be void when accessing an array, because of a concurrent change of the array. This can be an issue since a given index is not in a valid range anymore (e.g., testing `index < size`).

(ϱ_4) `add (get | index_of)`

The execution of `add` followed by the execution of `get` or `index_of` should be atomic. Otherwise, the added element needs no longer exist or its position in an array can be changed, when the program tries to obtain information about it.

2.4.2 Contracts for Concurrency with Parameters

The above definition of basic contracts is quite limited in some circumstances and can consider valid programs as erroneous (reports *false alarms*). Hence, in this section, there is defined an extension of basic contracts — *contracts with parameters* — which takes into consideration the data flow within function calls.

Example 2.4.2. Consider the following example from [11, 32], given Listing 2.1. There is a function `replace` that replaces item `a` in an array by item `b`. The implementation of this function comprises two atomicity violations:

(i) when `index_of` is invoked, item `a` does not need to be in the array anymore;

(ii) the acquired index can be obsolete when `set` is invoked.

A basic contract could cover this scenario by the clause ϱ_5:

(ϱ_5) `contains index_of set`

Nevertheless, this definition is too restrictive because the functions are required to be executed atomically only if `contains` and `index_of` have the same arguments `array` and `element`, `index_of` and `set` have the same argument `array`, and the returned value of `index_of` is used as the argument `index` of the function `set`.

```c
void replace(char *array, char a, char b)
{
    if (contains(array, a))
    {
        int index = index_of(array, a);
        set(array, index, b);
    }
}
```

Listing 2.1: An example of an atomicity violation with data dependencies [11]

In order to respect function call *parameters* and *return values* of functions in contracts, the basic contracts are extended by dependencies among functions in [11, 32] as follows. Function call parameters and return values are expressed as *meta-variables*. Further, if a contract should be required to be observed exclusively if the same object emerges as an argument or as the return value of multiple calls in a given call sequence, it may be denoted by using the same meta-variable at the position of all these occurrences of parameters and return values.

Clause ϱ_5 can be extended as follows (repeated application of meta-variables `X/Y/Z` requiring the same objects $\mathtt{o}_1/\mathtt{o}_2/\mathtt{o}_3$ to be used at the positions of `X/Y/Z`):

(ϱ'_5) `contains(X,Y) Z=index_of(X,Y) set(X,Z,_)`

The underscore indicates a *free meta-variable* that does not restrict the contract clause.

With the extension described above, it is possible to extend the contract from Section 2.4.1 as follows:

(ϱ'_1) `contains(X,Y) index_of(X,Y)`

(ϱ'_2) `Y=index_of(X,_) (get(X,Y) | set(X,Y,_) | remove(X,Y))`

2.4.3 Contracts for Concurrency with Spoilers

[14] [34] [23] [5] [1] [8] [9] [10] [7] [6] [2] [4] [36] [22] [28] [32] [27] [3] [24] [18] [30] [26] [31] [20] [17] [12]

Chapter 3

Atomer — Atomicity Violations Detector

This chapter describes the principles and limitations of the *static analyser Atomer* proposed as a module of *Facebook Infer* (introduced in Section 2.3) for finding some forms of *atomicity violations*. Atomer was proposed and in detail described in the bachelor's thesis of the author of this thesis [15]. Therefore, naturally, the following description in Section 3.2 is based on the mentioned thesis, and further, there is used information from [16].

Already existing solutions in this area (besides Atomer) discusses Section 3.1. In particular, it deals with other existing approaches and tools for finding atomicity violations, their advantages, disadvantages, features, availability, and so on.

In Section 3.3, there are discussed *limitations* and *shortcomings* of Atomer. Again, some of the thoughts mentioned in this section are taken into consideration in [15, 16].

3.1 Related Work

Atomer is slightly inspired by ideas from [11, 32]. In these papers, there is a proposal and implementation of a *static validation* for finding some forms of *atomicity violations* based on *grammars* and *parsing trees*. In the paper [11], there is also described and implemented a *dynamic* approach for the validation. The authors of these papers implemented a stand-alone prototype tool called *Gluon*[1] for analysing programs written in Java. It led to some promising experimental results, but the *scalability* of Gluon was still limited.[2] Moreover, Gluon is no more actively developed. This fact inspired the decision that eventually led to Atomer, namely, to get inspired by the ideas of [11, 32], but reimplement them in *Facebook Infer* redesigning it in accordance with the principles of Facebook Infer (described in Section 2.3), which should make the resulting tool more scalable. In the end, however, due to adapting the analysis for the context of Facebook Infer, the implementation of the analysis of Atomer is significantly different from [11, 32], as it is presented in Chapter 4

[1]**Gluon** is a tool for *static verification* of *contracts for concurrency* (see Section 2.4) in Java programs. It is available at `https://github.com/trxsys/gluon`.

[2]Some of the experiments performed by Gluon are also similarly performed by Atomer, which is mentioned in Sections 6.3 and 6.4.

of [15]. Furthermore, unlike [11, 32], the implementation aims at programs written in the *C language* using *PThread locks* to *synchronise concurrent threads*.

In Facebook Infer, there was already implemented an analysis called *Lock Consistency Violation*[3]. It is a part of *RacerD* [3, 14, 4]. This analysis finds atomicity violations in C++ and Objective C programs for reads/writes on single variables required to be executed atomically. Atomer is different; it finds atomicity violations for *sequences of functions* required to be executed atomically, i.e., it checks whether *contracts for concurrency* (see Section 2.4) hold. Moreover, it tries to automatically determine which sequences should indeed be executed atomically (hence, to *derive the contracts automatically*).

3.2 Design of Atomer and Its Principles

[[Říct, že Atomer není sound ani complete.]]

Atomer concentrates on checking *atomicity of execution of certain sequences of function calls*, which is often required for *concurrent programs'* correct behaviour. Atomer's principle is based on the assumption that sequences of function calls executed *atomically once* should probably be executed *always atomically*.

Listings in the below sections are pieces of exemplary programs written in C language (assuming *POSIX thread*, i.e., *PThread locks* and declared and initialised global variable `lock` of a type `phtread_mutex_t`).

As it has already been said, the proposal of Atomer is based on the concept of *contracts for concurrency* described in Section 2.4. In particular, the proposal considers *basic contracts* described in Section 2.4.1. Neither the contracts extended to *spoilers* and *targets* (from Section 2.4.3) nor contracts extended by *parameters* explained in Section 2.4.2 are so far taken into account.

In general, basic contracts for concurrency allow one to define *sequences of functions* required to be *executed atomically*, as explained in more detail in Section 2.4. Atomer is able to automatically derive candidates for such contracts and then verify whether the contracts are fulfilled. Both of these steps are done statically. The proposed analysis is divided into two parts (*phases of the analysis*):

Phase 1: Detection of (likely) *atomic sequences*.

Phase 2: Detection of *atomicity violations* (violations of the atomic sequences).

The phases are in-depth described in the sections below. Also, these phases of the analysis and its workflow are illustrated in Figure 3.1.

This section describes the proposal of Atomer in general. The concrete types of the *abstract states* (i.e. elements of the *abstract domain* $\boldsymbol{Q}$) and the *summaries* χ, along with the implementation of all necessary *abstract interpretation operators* are stated in Chapter 4 of [15]. However, actually, the abstract states $s \in \boldsymbol{Q}$ of both phases of the analysis are proposed as *sets*. So, in fact, the *ordering operator* $\sqsubseteq$ is implemented using testing for

[3]**Lock Consistency Violation** is an *atomicity violations* analysis implemented in *Facebook Infer*. It is described at `https://fbinfer.com/docs/checkers-bug-types/#lock-consistency-violation`.

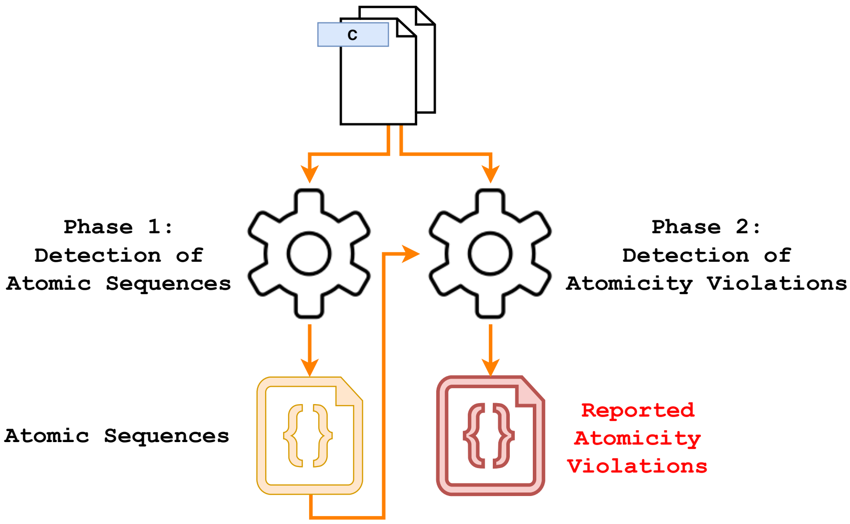


Figure 3.1: *Phases* of the Atomer's analysis and the analysis *high-level process* illustration

a *subset* (i.e. $s \sqsubseteq s' \Leftrightarrow s \subseteq s'$, where $s, s' \in \boldsymbol{Q}$), the *join operator* $\sqcup$ is implemented as the *set union* (i.e. $s \sqcup s' \Leftrightarrow s \cup s'$), and the *widening operator* $\triangledown$ is implemented using the join operator (i.e. $s \triangledown s' \Leftrightarrow s \sqcup s'$).

Function summaries are in below sections reduced to the output parts (*postconditions* R) only. The input parts of summaries (*preconditions* P) are in case of the proposed analysis always empty, because, so far, it is not necessary to have any preconditions for analysed functions. Thus, in this case, the Hoare triple *true* $\{Q\}$ R holds, where Q is an analysed program, i.e., $P = true$.

3.2.1 Phase 1: Detection of Atomic Sequences

Before detecting *atomicity violations* may begin, it is required to have *contracts* introduced in Section 2.4. **Phase 1** of Atomer is able to produce such contracts, i.e., it detects *sequences of functions* that should be *executed atomically*. Intuitively, the detection is based on looking for sequences of functions executed atomically on some path through a program. The assumption is that if it is *once needed to execute a sequence atomically*, it should probably be *always executed atomically*.

For a description of the analysis, it is first needed to introduce a notion of a *reduced sequence* of function calls. Such a sequence denotes a sequence in which the first appearance of each function is recorded only. The reason is to ensure the *finiteness* of the sequences derived by the analysis, and hence the analysis's *termination*. The detection of sequences of calls to be executed atomically is based on analysing all paths through the CFG of a function and generating all pairs $(A, B) \in \Sigma^* \times \Sigma^*$ (where Σ^* is a set of all possible sequences of functions from Σ from a given program) of reduced sequences of function calls for each path such that: Here, A is a reduced sequence of function calls that appear between the beginning of the function being analysed and the first lock, between an unlock and a subsequent lock, or between an unlock and the end of the function being analysed. B is a reduced sequence of function calls that follow the calls from A, and that appear between a lock and an unlock (or between a lock and the end of the function being analysed). Thus, the *abstract state*

$s \in \boldsymbol{Q}$ is defined as $2^{2^{\Sigma^* \times \Sigma^*}}$ (because there is a set of the (A, B) pairs for each program path).

It would be more precise to generate longer sequences of type $A_1 B_1 A_2 B_2 \ldots$, instead of the sets of the pairs (A, B). Nevertheless, it would be more challenging to ensure the above longer sequences' finiteness and the sets of these sequences' finiteness. Moreover, there would be a significantly larger *state space explosion problem* [33]. So, the proposed representation of the sets of pairs of sequences has been chosen to compromise accuracy and efficiency. However, the experiments described in Chapter 5 of [15] show that it needs even more pronounced abstraction for appropriate *scalability*.

Formally, the *initial abstract state* of a function is defined as $s_{init} = \{\{(\varepsilon, \varepsilon)\}\}$, where ε indicates an empty sequence. To formalise the analysis of a function, let `f` be a called leaf function. Further, let $s_{\mathtt{g}}$ be the abstract state of a function `g` being analysed before the function `f` is called. After the call of `f`, the abstract state will be changed as follows: $s_{\mathtt{g}} = \{p' \in 2^{\Sigma^* \times \Sigma^*} \mid \exists p \in s_{\mathtt{g}} : p' = \{(A', B') \in \Sigma^* \times \Sigma^* \mid \exists (A, B) \in p : (\neg actual(p, (A, B)) \wedge (A', B') = (A, B)) \vee [actual(p, (A, B)) \wedge [(lock \wedge (A', B') = (A, B \cdot \mathtt{f})) \vee (\neg lock \wedge (A', B') = (A \cdot \mathtt{f}, B))]]\}\}$, where *actual* is a Boolean function that determines whether a given (A, B) pair is the most recent pair of sequences of the current program state for a given program path. Furthermore, *lock* is a predicate indicating whether the current program state is inside an atomic block. Furthermore, let $s_{\mathtt{g}}$ be the abstract state of a function `g` being analysed before an unlock is called. After the unlock is called, a new (A, B) pair is created and labelled as an actual using the function *setActual* as follows: $s_{\mathtt{g}} = \{p' \in 2^{\Sigma^* \times \Sigma^*} \mid \exists p \in s_{\mathtt{g}} : p' = \{(A', B') \in \Sigma^* \times \Sigma^* \mid ((A', B') = (\varepsilon, \varepsilon) \wedge setActual(p, (A', B'))) \vee (A', B') \in p\}\}$.

Example 3.2.1. For an explanation of the computation of the sets of the pairs (A, B), assume that a state of the analysis of a program Q is the following sequence of function calls: `f, g`; and a state of the analysis of a program Q' is the following sequence of function calls: `f, g [m, n`. The square brackets are used to indicate an *atomic sequence* (the closing square bracket is missing in the case of the program Q', which means that the program state is currently inside an *atomic block*). The computed abstract state for the program Q is $s_Q = \{\{((\mathtt{f}, \mathtt{g}), \varepsilon)\}\}$, and for the program Q', it is $s_{Q'} = \{\{((\mathtt{f}, \mathtt{g}), (\mathtt{m}, \mathtt{n}))\}\}$. Now, if the next instruction is a call of a function `x`, in the case of the program Q, the call will be added to the first A sequence, and in the case of the program Q', the call will be added to the first B sequence as follows: $s_Q = \{\{((\mathtt{f}, \mathtt{g}, \mathtt{x}), \varepsilon)\}\}$, $s_{Q'} = \{\{((\mathtt{f}, \mathtt{g}), (\mathtt{m}, \mathtt{n}, \mathtt{x}))\}\}$. Subsequently, if the next step in the program Q is a lock call, the next function calls will be added to the first B sequence of the set s_Q. And if the next step in the program Q' is an unlock call, it will be created a new element of the first set of the set $s_{Q'}$, and the next function calls will be added to the A sequence of this element. Finally, if a function `y` is called, the resulting sets will look like follows: $s_Q = \{\{((\mathtt{f}, \mathtt{g}, \mathtt{x}), (\mathtt{y}))\}\}$, $s_{Q'} = \{\{((\mathtt{f}, \mathtt{g}), (\mathtt{m}, \mathtt{n}, \mathtt{x})), ((\mathtt{y}), \ \varepsilon)\}\}$. Note that the final sequences of function calls look like follows: `f, g, x [y` and `f, g [m, n, x] y` for the programs Q, and Q', respectively.

The *summary* $\chi_{\mathtt{f}} \in 2^{\Sigma^*} \times \Sigma^*$ of a function `f` is then defined as $\chi_{\mathtt{f}} = (\boldsymbol{B}, AB)$, where:

- $\boldsymbol{B}$ is a set constructed that contains all the B sequences that appear on program paths through `f`, i.e., $\boldsymbol{B} = \{B' \in \Sigma^* \mid \exists p \in s_{\mathtt{f}} : \exists (A, B) \in p : B \neq \varepsilon \wedge B' = B\}$, where $s_{\mathtt{f}}$ is the abstract state at the end of an interpretation of `f`. In other words, this component of the summary is a set of sequences of atomic function calls appearing in an analysed function.

- AB is a *concatenation* of all the A and B sequences with removed duplicates of function calls. In particular, assume that there is the following computed set of (A, B) pairs: $\{(A_1, B_1), (A_2, B_2), \ldots, (A_n, B_n)\}$, then the result is concatenated sequence $A_1 \cdot B_1 \cdot A_2 \cdot B_2 \cdot \ldots \cdot A_n \cdot B_n$ with removed duplicates. Formally:

$$AB = \mathit{reduce}\Big(\dot{\bigcup_{ab \in AB'}} ab\Big)$$

  where $AB' = \{ab \in \Sigma^* \mid \exists p \in s_{\texttt{f}} : \exists (A, B) \in \underset{\cdot}{p} : ab = A \cdot B\}$, *reduce* is a function that removes duplicates of function calls, and $\bigcup$ concatenates all sequences of a set. Intuitively, in this component of the summary, it is gathered occurrences of all called functions within an analysed function obtained by concatenation of all the A and B sequences.

AB is recorded to analyse functions higher in the *call hierarchy* since locks/unlocks can appear in such a *higher-level* function.

Example 3.2.2. For instance, the analysis of the function `f` from Listing 3.1 produces the following sequences:

$$\overbrace{\texttt{a},\ \cancel{\texttt{a}}}^{A_1}\ \overbrace{[\texttt{a},\ \cancel{\texttt{a}},\ \texttt{b}]}^{B_1}\ \overbrace{\texttt{a},\ \cancel{\texttt{a}}}^{A_2}\ \overbrace{[\texttt{a},\ \texttt{c}]}^{B_2}\ \overbrace{\cancel{\texttt{a}}}^{\cancel{A_3}}\ \overbrace{\cancel{[\texttt{a},\ \texttt{c},\ \texttt{c}]}}^{\cancel{B_3}}$$

The functions `a, b, c` are not deeper analysed because it is assumed that these functions are leaf nodes of the *call graph*. The strikethrough of the functions `a` and `c` denotes removing already recorded function calls in the A and B sequences. The strikethrough of the entire sequence `a [a, c, c]` means discarding sequence already seen before. For the above, the abstract state at the end of an interpretation of the function `f` is $s_{\texttt{f}} = \{\{((\texttt{a}), (\texttt{a}, \texttt{b})), ((\texttt{a}), (\texttt{a}, \texttt{c})), (\varepsilon, \varepsilon)\}\}$. The derived summary $\chi_{\texttt{f}}$ for the function `f` is $\chi_{\texttt{f}} = (\boldsymbol{B}, AB)$, where:

- $\boldsymbol{B} = \{(\texttt{a}, \texttt{b}), (\texttt{a}, \texttt{c})\}$, i.e., B_1 and B_2;
- $AB = (\texttt{a}, \texttt{b}, \texttt{c})$, i.e., the concatenation of A_1, B_1, A_2, and B_2 from which duplicate function calls were removed.

Further, it is demonstrated how the results of the analysis of *nested functions* are used during the detection of atomic sequences. The result of the analysis of a nested function is used as follows. When calling an already analysed function, one plugs the sequences from the second component of its summary (i.e. the AB sequence) into the most recent A or B sequence of all the program paths (where a program path corresponds to a single element of an abstract state, i.e., a set of the (A, B) pairs). In particular, assume that (A, B) is the most recent pair of sequences of the program state of a path being analysed. Subsequently, it is called a function `f` with a *non-empty summary* (i.e. $AB \neq \varepsilon$). If the current program state of an analysed function is inside an atomic block, the analysis in this step will transform the pair (A, B) to a new (A', B') pair as follows: $(A', B') = (A, B \cdot \texttt{f} \cdot AB)$. Otherwise, $(A', B') = (A \cdot \texttt{f} \cdot AB, B)$. In such cases where a summary is empty, i.e., there are no function calls in a called function, or it is a leaf node of the call graph, it is appended just the function name to the most recent A or B sequences of all

```c
void f()
{
    a(); a();

    pthread_mutex_lock(&lock); // (a, b)
    a(); a(); b();
    pthread_mutex_unlock(&lock);

    a(); a();

    pthread_mutex_lock(&lock); // (a, c)
    a(); c();
    pthread_mutex_unlock(&lock);

    a();

    pthread_mutex_lock(&lock); // (a, c)
    a(); c(); c();
    pthread_mutex_unlock(&lock);
}
```

Listing 3.1: A code snippet used for an illustration of the derivation of *sequences of functions called atomically*

the program paths. To formalise this process, let `f` be a called function that was already analysed, and the second component of its summary is AB. Further, let $s_{\mathtt{g}}$ be the abstract state of a function `g` being analysed before the function `f` is called. After the call of `f`, the abstract state will be changed as follows: $s_{\mathtt{g}} = \{p' \in 2^{\Sigma^* \times \Sigma^*} \mid \exists\, p \in s_{\mathtt{g}} : p' = \{(A', B') \in \Sigma^* \times \Sigma^* \mid \exists\,(A, B) \in p : (\neg actual(p, (A, B)) \wedge (A', B') = (A, B)) \vee [actual(p, (A, B)) \wedge [(lock \wedge (A', B') = (A, B \cdot \mathtt{f} \cdot AB)) \vee (\neg lock \wedge (A', B') = (A \cdot \mathtt{f} \cdot AB, B))]]\}\}$.

Example 3.2.3. This example shows how the function `g` from Listing 3.2 would be analysed using the result of the analysis of the function `f` from Listing 3.1. So the analysis of the function `g` produces the following sequence:

$$\mathtt{a,\ f,\ \cancel{a},\ b,\ c\ [f,\ a,\ b,\ c]}$$

For the above, the abstract state at the end of an interpretation of the function `g` is $s_{\mathtt{g}} = \{\{((\mathtt{a,\ f,\ b,\ c}), (\mathtt{f,\ a,\ b,\ c})), (\varepsilon, \varepsilon)\}\}$. The derived summary $\chi_{\mathtt{g}}$ for the function `g` is $\chi_{\mathtt{g}} = (\{(\mathtt{f,\ a,\ b,\ c})\}, (\mathtt{a,\ f,\ b,\ c}))$.

Cases Where Lock/Unlock Calls Are Not Paired in a Function For treating cases where *lock/unlock calls are not paired* in a function — as demonstrated in Listing 3.3 — in Atomer, the following solution is implemented.

Everything is unlocked at the end of a function, i.e., one *virtually* appends an unlock to the end of the function if it is necessary. Then, for the function `x` from Listing 3.3, the atomic section is virtually closed. Hence, there is detected an atomic sequence (`a`). In particular, the summary is as follows: $\chi_{\mathtt{x}} = (\{(\mathtt{a})\}, (\mathtt{a}))$.

```c
void g()
{
    a(); f();

    pthread_mutex_lock(&lock); // (f, a, b, c)
    f();
    pthread_mutex_unlock(&lock);
}
```

Listing 3.2: A code snippet used to illustrate the derivation of *sequences of functions called atomically* with a *nested function calls* (function `f` is defined in Listing 3.1)

Moreover, *all unlock calls not preceded by a lock are ignored.* Thus, in the function `y` from Listing 3.3, there are not detected any atomic sequences: $\chi_{\mathtt{y}} = (\emptyset, (\mathtt{a}))$.

```c
void x()
{
    pthread_mutex_lock(&lock); // (a)
    a();
}
void y()
{
    a();
    pthread_mutex_unlock(&lock);
}
```

Listing 3.3: A code snippet used to illustrate treating cases where *lock/unlock calls are not paired* in a function

Summary of the Phase 1 The above detection of atomic sequences was implemented and also validated on a set of sample programs created for testing purposes, which is described in [15]. The derived sequences of calls assumed to execute atomically — the $\boldsymbol{B}$ sequences — from the summaries of all analysed functions are stored into a file used during **Phase 2**, which is described below.

3.2.2 Phase 2: Detection of Atomicity Violations

In the second phase of the analysis, i.e., when *detecting violations* of the atomic sequences obtained from **Phase 1**, the analysis looks for *pairs of functions* that *should be called atomically* (or just for single functions if there is only one function call in an atomic sequence) while this is not the case on some path through the CFG. The pairs of function calls to be checked for atomicity are obtained as follows: For each function `f` with the *summary* $\chi_{\mathtt{f}} = (\boldsymbol{B}, AB)$ in a given program Q, it is taken the first component $\boldsymbol{B}$ of the summary $\chi_{\mathtt{f}}$, i.e., $\boldsymbol{B} = \{B_1, B_2, \ldots, B_n\}$, and it is taken *every pair* $(\mathtt{x}, \mathtt{y}) \in \Sigma \times \Sigma$ of functions that appear as a *substring* in some of the B_i sequences, i.e., $B_i = w \cdot \mathtt{x} \cdot \mathtt{y} \cdot w'$ for some sequences w, w'. Note that `x` could be ε (an empty sequence) if some B_i consists of a single function. All these "atomic pairs" are put into the set $\Omega \in 2^{\Sigma \times \Sigma}$. More formally, $\Omega = \{(\mathtt{x}, \mathtt{y}) \in \Sigma \times \Sigma \mid \exists (\boldsymbol{B}, AB) \in X_Q : \exists B \in \boldsymbol{B} : (|B| = 1 \wedge (\mathtt{x}, \mathtt{y}) = (\varepsilon, B)) \vee (|B| >$

$1 \wedge \exists\, w, w' \in \Sigma^* : B = w \cdot \mathtt{x} \cdot \mathtt{y} \cdot w' \wedge (\mathtt{x}, \mathtt{y}) \neq (\varepsilon, \varepsilon))\}$, where Σ^* is a set of all possible sequences of functions from Σ from a given program, and $X_Q \in 2^{2^{\Sigma^*} \times \Sigma^*}$ is a set of all summaries of the program Q.

Example 3.2.4. For instance, assume that in **Phase 1**, there was analysed a function `f`. Which produced the summary $\chi_{\mathtt{f}} = (\boldsymbol{B}, AB)$, where $\boldsymbol{B} = \{(\mathtt{a}, \mathtt{b}, \mathtt{c}), (\mathtt{a}, \mathtt{c}, \mathtt{d})\}$, i.e., a set of sequences of functions that should be called atomically. The analysis will then look for the following pairs of functions that are not called atomically: $\Omega = \{(\mathtt{a}, \mathtt{b}), (\mathtt{b}, \mathtt{c}), (\mathtt{a}, \mathtt{c}), (\mathtt{c}, \mathtt{d})\}$.

An element of this phase's *abstract state* is a triple $(\mathtt{x}, \mathtt{y}, \Delta) \in \Sigma \times \Sigma \times 2^{\Sigma \times \Sigma}$, where $(\mathtt{x}, \mathtt{y})$ is a pair of the most recent function calls, and Δ is a *set of pairs that violate atomicity*. Thus, the abstract state $s \in \boldsymbol{Q}$ is defined as $2^{\Sigma \times \Sigma \times 2^{\Sigma \times \Sigma}}$. Whenever a function `f` is called, it is created a new pair $(\mathtt{x}', \mathtt{y}')$ of the most recent function calls from the previous pair $(\mathtt{x}, \mathtt{y})$ (i.e. $(\mathtt{x}', \mathtt{y}') = (\mathtt{y}, \mathtt{f})$). Further, when the current program state is not inside an atomic block, it is checked whether the new pair (or just the last call) violates atomicity (i.e. $(\mathtt{x}', \mathtt{y}') \in \Omega \vee (\varepsilon, \mathtt{y}') \in \Omega$). When it does, it is added to the set Δ of pairs that violate atomicity.

Formally, the *initial abstract state* of a function is defined as $s_{init} = \{(\varepsilon, \varepsilon, \emptyset)\}$. To formalise the analysis of a function, let `f` be a called leaf function. Further, let $s_{\mathtt{g}}$ be the abstract state of a function `g` being analysed before the function `f` is called. After the call of `f`, the abstract state will be changed as follows: $s_{\mathtt{g}} = \{(\mathtt{x}', \mathtt{y}', \Delta') \in \Sigma \times \Sigma \times 2^{\Sigma \times \Sigma} \mid \exists\, (\mathtt{x}, \mathtt{y}, \Delta) \in s_{\mathtt{g}} : (\mathtt{x}', \mathtt{y}') = (\mathtt{y}, \mathtt{f}) \wedge [(\neg lock \wedge \Delta' = \{(\mathtt{x}'', \mathtt{y}'') \in \Sigma \times \Sigma \mid (\mathtt{x}'', \mathtt{y}'') \in \Delta \vee [((\mathtt{x}'', \mathtt{y}'') = (\mathtt{x}', \mathtt{y}') \vee (\mathtt{x}'', \mathtt{y}'') = (\varepsilon, \mathtt{y}')) \wedge (\mathtt{x}'', \mathtt{y}'') \in \Omega]\}) \vee (lock \wedge \Delta' = \Delta)]\}$.

The analysis of functions with *nested function calls* and cases where *lock/unlock calls not not paired* in functions are handled analogically as in **Phase 1**. For detailed examples, see [15].

Example 3.2.5. To demonstrate the detection of an atomicity violation, assume the functions `f` and `g` from Listing 3.4. The set of atomic sequences of the function `f` with the summary $\chi_{\mathtt{f}} = (\boldsymbol{B}, AB)$ is $\boldsymbol{B} = \{(\mathtt{b}, \mathtt{c})\}$, thus $\Omega = \{(\mathtt{b}, \mathtt{c})\}$. In the function `g`, an atomicity violation is detected because the pair of functions `b` and `c` is not called atomically (under a lock).

Summary of the Phase 2 Like in the first phase of the analysis, **Phase 2** was implemented and validated on a set of purposeful sample programs, as it is described in [15]. The sets of atomicity violations Δ from individual functions are the final reported atomicity violations seen by a user.

3.3 Atomer's Limitations

Atomer was proposed as it is detailed in Section 3.2. The analyser was implemented, and it is working as expected. Moreover, it can be used in practice to analyse various kinds of programs, and it may find *real atomicity related bugs*. Nevertheless, there are still several limitations and cases where Atomer would not work correctly, i.e., cases not considered during the original proposal. Some of these cases were briefly discussed in [15], and further described in [16].

```c
void f()
{
    a();

    pthread_mutex_lock(&lock); // (b, c)
    b(); c();
    pthread_mutex_unlock(&lock);

    d();
}
void g()
{
    a(); b(); c(); d(); // ATOMICITY_VIOLATION: (b, c)
}
```

Listing 3.4: Example of an *atomicity violation*

So far, Atomer does not work with *nested locks*, i.e., it does not distinguish *different locks* used in a program. Only calls of locks/unlocks are identified, and parameters of these calls (*lock objects*) are not considered. So, if there are several lock objects used, the analysis does not work correctly. Although this may happen in *real-life programs*, insomuch as one could have another (smaller) atomic section inside a current atomic section (this does not have to be evident at first because the *inner atomic section* could be, e.g., included via a macro). For instance: `lock(A); lock(B); ... unlock(B); unlock(A);`. Another possibility is an *alternating sequence of locks*, e.g., two locks are locked at first, and then, they are unlocked in the same order, i.e., `lock(A); lock(B); ... unlock(A); unlock(B);`.

Atomer considers only *basic contracts for concurrency*, which are defined in Section 2.4.1. It is quite limited in some circumstances and therefore, Atomer can report *false alarms*. The basic contracts do not take into consideration the *data flow* within function calls. However, actually, a better idea is to work with the assumption that a sequence of function calls must be atomic only if they *handle the same data*. Assume that there are functions `f, g` manipulating with the same container `c` as follows: `f(c); g(c);`. These are called atomically. Somewhere else — where `f, g` are not called atomically — it does not necessarily mean atomicity violation because they can be invoked with different arguments, which could be valid. This behaviour corresponds to the *extended contracts with parameters* (see Section 2.4.2). Another, a more complex limitation is that basic contracts do not consider any *contextual information*. It would be more precise to consider as atomicity violations such sequences that could be violated only by particular ("dangerous") function calls, not by any calls. For example, assume that there is the following sequence of functions called atomically: `f(); g();`. While somewhere else, these functions are not called atomically, it does not necessarily mean that it is an atomicity violation because, in this particular context, non of the "dangerous" functions can be executed by any concurrent thread. The *extended contracts with spoilers* formally describe these cases in Section 2.4.3.

Another limitation of Atomer is that it supports only the analysis of programs written in the *C language* that uses *PThread* locks to *synchronise concurrent threads*. Naturally, in practice, many other *types of locks* for synchronisation of concurrent threads or even *synchronisation of concurrent processes* are used. Although the first version of Atomer can analyse C programs with other types of locks, these locks are not recognised as locks. Thus, the analysis would not work as expected. Of course, it would be useful also to analyse other languages than just C. As described in Section 2.3; *Facebook Infer* is capable of analysing programs written in C, C++, Objective C, and Java (and C#). An analysis algorithm could then be the same for all these languages because the Infer's *intermediate language* is analysed, instead of directly analysing the input languages. Again, Atomer should be able to analyse the above languages, but it was not tested in [15]. However, most importantly, other languages might use *very different locks types*, and these would not be recognised.

One of the main reasons that Atomer reports *false alarms* is that in *critical sections*, in practice, there are sometimes called *generic functions* that do not influence atomicity violations (such as functions for printing to the standard output, functions for recasting variable to different types, functions related to iterators, and whatever other "safe" functions for particular program types). Often, to find some atomicity violations, it is sufficient to focus only on certain "critical" functions. In practice, another issue is that in an analysed program, there could be "large" critical sections or critical sections in which appear function calls with a *deep hierarchy of nested function calls*. All the above cases could cause massive and "imprecise" atomic sequences that are the source of false alarms. However, regardless of the above issues, Atomer can still report quite some false alarms. It is due to the assumption that *sequences called atomically once* should *always be called atomically*, but this does not always have to hold. None of the above reasons that could generate false alarms is resolved in the first version of Atomer.

A remarkable problem (though it is not directly a problem of Atomer) is identifying whether a reported atomicity violation is a *real bug* or whether it is just a false alarm. It could be really challenging, especially in *extensive real-life* programs.

Furthermore, Atomer does not consider a locking using *trylock* functions, i.e., functions equivalent to lock functions, except that if the lock object is currently locked, the call of the trylock shall return immediately, i.e., no waiting. It can be determined from the return value of the trylock whether the lock succeeded or not. These types of locks are used (though not so often) in practice as well. In Atomer, trylocks are so far not identified as locks. The question is how to propose an extension of Atomer that would opportunely handle trylocks.

Regarding the *scalability*, Atomer can have problems with more *extensive* and *complex* programs (problems with the *memory* as well as problems with the *analysis time*). A problem is working with the sets of (A, B) pairs of *sequences* in abstract states, and working with *sequences* of atomic calls in summaries. It may be necessary to store many of these sequences, and they could be very long (due to all different paths through the CFG of an analysed program). This leads to the *state space explosion problem* [33].

Solutions for some of the above problems and limitations were proposed in Chapter 4 and further implemented in a new version of Atomer, detailed in Chapter 5.

Chapter 4

Proposal of Enhancements for Atomer

This chapter describes the proposed solutions for Atomer's limitations stated in Section 3.3, i.e., solutions that enhance *precision* and *scalability* of the analysis performed by Atomer. To formally define these enhancements, notions and symbols introduced in Section 3.2 are used. Some of the enhancements were described already in [16].

In the following sections, to give an intuition, there are used listings with C programs that use *PThread locks* and assume declared and initialised global variables `lock`, `lockA`, `lockB`, ... (of a type `phtread_mutex_lock`).

Section 4.1 proposes an optimisation of Atomer's scalability. The following sections 4.2, 4.3, and 4.4 covers precision improvements, i.e., an extension of Atomer by additional features that improve its ability to cope with cases that were not supported in the first version of Atomer, and that can be seen in *real-life code*. A description of the implementation of all the below improvements is available in Chapter 5.

In the description in the below Sections, the enhancements stated in the preceding Sections of a given Section are considered.

4.1 Approximation of Sequences by Sets

Because Atomer can have *scalability* problems when analysing more *extensive* and *complex* programs (problems with the *memory* as well as problems with the *analysis time*), it was proposed the following optimisation. It seems promising to *approximate* (*abstraction refinement*) working with the sets of (A, B) pairs of *sequences* of function calls in *abstract states* (during **Phase 1**) by working with the sets of (A, B) pairs of **sets** of function calls. Elements of these pairs are also occurring in *summaries* of the first phase, and they are used during **Phase 2**. Thus, it is needed to make a certain approximation in these structures and algorithms likewise. The approximated phases of the analysis and its collaboration are illustrated in Figure 4.1 (one can compare that with the illustration of the first version of Atomer in Figure 3.1).

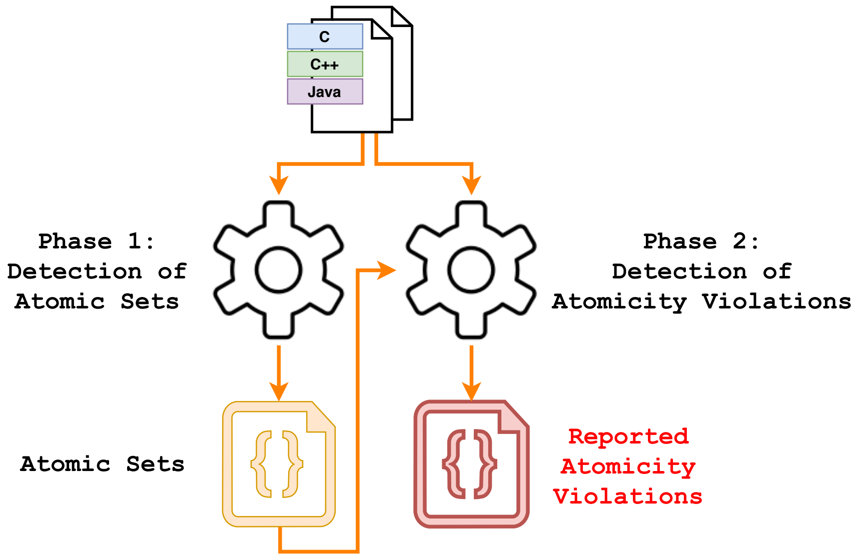


Figure 4.1: An illustration of the *phases* of the Atomer's analysis and the *high-level analysis process* with an *approximation* of working with *sequences* by working with *sets* (moreover, it can be visible that Atomer now accepts programs also written in *C++* and *Java* languages, which is described in Section 5.3)

In particular, this proposed solution is more scalable because the order of stored function calls is not relevant anymore while working with sets. Therefore, less memory is required because the same sets of function calls are not stored multiple times. The analysis is also faster since there are stored fewer sets of function calls to work with. On the other hand, the analysis is less accurate because the new approach causes some loss of information. In practice, this loss of information could eventually lead to *false alarms*. However, the number of such false alarms should not often be so significant. Moreover, later, there are presented some techniques that try to rid of these false alarms.

4.1.1 Approximation of the Abstract State and the Summary in Phase 1

The *detection of sequences of calls to be executed atomically* is now based on analysing all paths through the CFG of a function and generating all pairs $(A, B) \in 2^{\Sigma} \times 2^{\Sigma}$ (where Σ is a set of all function names in a given program) of **sets** of function calls for each path. Here, A, B are not *reduced sequences* (the notion of a reduced sequence is not needed anymore), but sets, and their semantics is preserved. So, the *abstract state* $s \in \boldsymbol{Q}$ is redefined as $2^{2^{2^{\Sigma}} \times 2^{\Sigma}}$.

Further, in all the defined algorithms and definitions, it is sufficient to work with:

- *sets* of functions 2^{Σ}, instead of *sequences* of functions Σ^*;
- *empty sets* $\emptyset$, instead of *empty sequences* ε;
- and *union* of sets $\cup$, instead of a *concatenation* of sequences $\cdot$.

The above implies that the *initial abstract state* of a function is changed to $s_{init} = \{\{(\emptyset, \emptyset)\}\}$. During the analysis of a function g with an abstract state $s_{\mathtt{g}}$, when a leaf function f is called, the abstract state's transformation is changed as follows: $s_{\mathtt{g}} = \{p' \in 2^{2^{\Sigma} \times 2^{\Sigma}} \mid \exists\, p \in s_{\mathtt{g}} : p' = \{(A', B') \in 2^{\Sigma} \times 2^{\Sigma} \mid \exists\,(A, B) \in p : (\neg actual(p, (A, B)) \wedge (A', B') = (A, B)) \vee [actual(p, (A, B)) \wedge [(lock \wedge (A', B') = (A, B \cup \{\mathtt{f}\})) \vee (\neg lock \wedge (A', B') = (A \cup \{\mathtt{f}\}, B))]]\}\}$. Further, when an unlock is called, a new (A, B) pair is created as follows: $s_{\mathtt{g}} = \{p' \in 2^{2^{\Sigma} \times 2^{\Sigma}} \mid \exists\, p \in s_{\mathtt{g}} : p' = \{(A', B') \in 2^{\Sigma} \times 2^{\Sigma} \mid ((A', B') = (\emptyset, \emptyset) \wedge setActual(p, (A', B'))) \vee (A', B') \in p\}\}$. Other definitions (e.g. calling an already analysed *nested* function) will be modified analogically.

Another approximation was made in *summaries*. The first component of the summary has to be changed to a set of sets of function calls because it is constructed from the B items from abstract states, which are now sets. The second component of the summary can be changed to a set of function calls, because even before, it was a reduced sequence of all the (A, B) pairs. Therefore, the order of function calls was significantly approximated even so. Moreover, it is used to analyse functions higher in the *call hierarchy* where it is appended to A or B, which are now sets. Thus, it would make no sense to store it in summaries as a sequence. Formally, the summary $\chi_{\mathtt{f}} \in 2^{2^{\Sigma}} \times 2^{\Sigma}$ of a function f is redefined as $\chi_{\mathtt{f}} = (\boldsymbol{B}, AB)$, where:

- $\boldsymbol{B} = \{B' \in 2^{\Sigma} \mid \exists\, p \in s_{\mathtt{f}} : \exists\,(A, B) \in p : B \neq \emptyset \wedge B' = B\}$, where $s_{\mathtt{f}}$ is the abstract state at the end of an interpretation of f.
- $AB = \bigcup\limits_{ab \in AB'} ab$, where $AB' = \{ab \in 2^{\Sigma} \mid \exists\, p \in s_{\mathtt{f}} : \exists\,(A, B) \in p : ab = A \cup B\}$.

Example 4.1.1. For demonstrating the approximation of the analysis with sets, assume functions f and g from Listing 4.1. Further assume, that a, b, x, y are leaf nodes of the *call graph*. Before the approximation, when the analysis was working with sequences of function calls, **Phase 1** of the analysis produced the following abstract states and summaries while analysing the functions:

- $s_{\mathtt{f}} = \{\{((\mathtt{a}, \mathtt{b}), (\mathtt{x}, \mathtt{y})), ((\mathtt{b}, \mathtt{a}), (\mathtt{y}, \mathtt{x})), (\varepsilon, \varepsilon)\}\}$, $\chi_{\mathtt{f}} = (\{(\mathtt{x}, \mathtt{y}), (\mathtt{y}, \mathtt{x})\}, (\mathtt{a}, \mathtt{b}, \mathtt{x}, \mathtt{y}))$;
- $s_{\mathtt{g}} = \{\{((\mathtt{b}, \mathtt{a}), (\mathtt{y}, \mathtt{x})), (\varepsilon, \varepsilon)\}\}$, $\chi_{\mathtt{g}} = (\{(\mathtt{y}, \mathtt{x})\}, (\mathtt{b}, \mathtt{a}, \mathtt{y}, \mathtt{x}))$.

Whereas, after the approximation, the produced abstract states and summaries are as follows (i.e. there is only one "atomic set", and the summaries and abstract states are the same for both functions because there are the same locked/unlocked function calls, only the order of calls is different):

- $s_{\mathtt{f}} = \{\{(\{\mathtt{a}, \mathtt{b}\}, \{\mathtt{x}, \mathtt{y}\}), (\emptyset, \emptyset)\}\}$, $\chi_{\mathtt{f}} = (\{\{\mathtt{x}, \mathtt{y}\}\}, \{\mathtt{a}, \mathtt{b}, \mathtt{x}, \mathtt{y}\})$;
- $s_{\mathtt{g}} = \{\{(\{\mathtt{a}, \mathtt{b}\}, \{\mathtt{x}, \mathtt{y}\}), (\emptyset, \emptyset)\}\}$, $\chi_{\mathtt{g}} = (\{\{\mathtt{x}, \mathtt{y}\}\}, \{\mathtt{a}, \mathtt{b}, \mathtt{x}, \mathtt{y}\})$.

4.1.2 Approximation with Sets in Phase 2

Detecting violations of atomicity works almost the same way as before the approximation. There is only one difference. Before the approximation, it was detected violations of atomic

```
void f()
{
    a(); b();

    pthread_mutex_lock(&lock); // (x, y) -> {x, y}
    x(); y();
    pthread_mutex_unlock(&lock);

    b(); a();

    pthread_mutex_lock(&lock); // (y, x) -> {x, y}
    y(); x();
    pthread_mutex_unlock(&lock);
}
void g()
{
    b(); a();

    pthread_mutex_lock(&lock); // (y, x) -> {x, y}
    y(); x();
    pthread_mutex_unlock(&lock);
}
```

Listing 4.1: A code snippet used to illustrate the Atomer's **Phase 1** *approximation* of the analysis with *sets of function calls*

sequences obtained from **Phase 1**. Now, **atomic sets** are obtained; hence, detection of violations of atomic sets is performed. Again, the analysis looks for *pairs of functions that should be called atomically* while this is not the case on some path through the CFG. This algorithm is identical to the algorithm before the approximation.

Nevertheless, it is needed to propose a new algorithm that derives the pairs of function calls (from the atomic sets) to be checked for atomicity (i.e. the set $\Omega \in 2^{\Sigma \times \Sigma}$). In order to obtain the pairs, it is taken a union of sets that contain all 2-element *variations* of single atomic sets (i.e. all the possible pairs). Formally, let Q be an analysed program, and let $X_Q \in 2^{2^{2^{\Sigma}} \times 2^{\Sigma}}$ be a set of all *summaries* of the program Q. Then, all the atomic pairs (the first item of a pair may be empty if an atomic set consists of a single function) are obtained as follows: $\Omega = \{(\mathtt{x}, \mathtt{y}) \in \Sigma \times \Sigma \mid \exists\, (\boldsymbol{B}, AB) \in X_Q : \exists\, B \in \boldsymbol{B} : (|B| = 1 \wedge (\mathtt{x}, \mathtt{y}) \in \{\varepsilon\} \times B) \vee (|B| > 1 \wedge (\mathtt{x}, \mathtt{y}) \in B \times B \wedge \mathtt{x} \neq \mathtt{y})\}$.

Example 4.1.2. For example, assume that in **Phase 1**, there was analysed a function `f`, which produced the summary $\chi_f = (\boldsymbol{B}, AB)$. Assume that before the approximation, a set of sequences of functions that should be called atomically was as follows: $\boldsymbol{B} = \{(\mathtt{a}, \mathtt{b}, \mathtt{c})\}$. Then, the analysis looked for the following pairs of functions that are not called atomically: $\Omega = \{(\mathtt{a}, \mathtt{b}), (\mathtt{b}, \mathtt{c})\}$. Since the result of the first component of the summary was changed to the following set of sets: $\boldsymbol{B} = \{\{\mathtt{a}, \mathtt{b}, \mathtt{c}\}\}$, the analysis now looks for the following pairs of functions that are not called atomically (all 2-element variations): $\Omega = \{(\mathtt{a}, \mathtt{b}), (\mathtt{a}, \mathtt{c}), (\mathtt{b}, \mathtt{a}), (\mathtt{b}, \mathtt{c}), (\mathtt{c}, \mathtt{a}), (\mathtt{c}, \mathtt{b})\}$.

4.2 Advanced Manipulation with Locks

In the first version of Atomer, *different locks* are not distinguished at all. Only calls of locks/unlocks are identified, and parameters of these calls (*lock objects*) are not considered. In order to consider lock objects, it was proposed distinguishing between them using Facebook Infer's built-in mechanism called *access paths*, explained in Section 4.2.1. The analyser does not perform a general *alias analysis*, i.e., it is not performed a precise analysis for saying when arbitrary pairs of accesses to lock objects may alias. During the analysis (both phases), each *atomic section* is identified by an access path of a lock that guards the section, see Sections 4.2.2, 4.2.3. Because *syntactically identical access paths* are used as the intuition for distinguishing atomic sections, some *atomicity violations* could be missed (or some *false alarms* could be reported) due to distinct access paths that refer to the same memory. However, it vastly simplifies the analysis, and the stress is put on finding likely violations.

4.2.1 Access Paths

The *syntactic access paths* [21] represent *heap locations* via the paths used to access them, i.e., a base variable followed by a sequence of fields. More formally, let Var be a set of all variables that can occur in a given program. Let $Field$ be a set of all possible field names that can be used in a given program (e.g. structure fields). Then, an access path π from the set Π of all access paths is defined as follows:

$$\pi \in \Pi ::= Var \times Field^*$$

Access paths are already implemented in Facebook Infer. For instance, the principle of using access paths is used in an existing analyser in Facebook Infer — RacerD [3] — for data race detection. In general, no sufficiently precise *alias analysis* works *compositionally* and at *scale*. That is the motivation for using access paths in Facebook Infer.

Given a pair of accesses to lock objects, to determine whether these locks are equal, it is needed to answer the following question: "Can the accesses touch the same address?". Remarkably, according to the authors of [3], access paths alone *almost* convey enough semantic information to answer the above question on their own. If two access paths are syntactically equal, it is almost (but not quite) true that they must refer to the same address. Syntactically identical paths can refer to different addresses if (i) they refer to different instances of the same object or (ii) a prefix of the path is reassigned along one execution trace, but not the other. These conditions cannot hold if an access path is *stable*, i.e., if none of its proper prefixes appears in assignments during a given execution trace, then it touches the same memory as all other stable accesses to the syntactic path. So, access paths' syntactic equality is a reasonably efficient way to say (in an *under-approximate fashion*) that heap access touches the same address. Also, by using access paths, RacerD detected many errors in real-world programs, proving that the use of access paths can reveal real errors. This is why it was decided to use this principle to represent locks in Atomer.

4.2.2 Distinguishing Multiple Locks in Phase 1

The *detection of sets of calls to be executed atomically* is based on generating all pairs $(A, B) \in 2^\Sigma \times 2^\Sigma$. Now, it is needed to store *access paths* of locks that guard calls executed atomically, i.e., the B sets. Therefore, these pairs are extended to the triples $(A, B, \pi) \in 2^\Sigma \times 2^\Sigma \times \Pi$, where the third component is an access path that identifies a *lock object* which locks an atomic section that contains the calls from B. Note, that π could also be ε (i.e. $\varepsilon \in \Pi$), which is a special case when there is no lock associated to the (A, B) pair so far, i.e., B is empty as well, and a lock was not called yet. The *abstract state* $s \in \boldsymbol{Q}$ is now defined as $2^{2^{2^\Sigma \times 2^\Sigma \times \Pi}}$. When a function is called, it is appended to the A set of the triple where $\pi = \varepsilon$, i.e., the triple without an associated lock. Also, it is appended to all the triples that have some lock which is currently locked. When a lock is called, its identifier is associated to the triple without any lock associated to it (which is then labelled as the currently locked lock), and it is created a new triple without a lock. Finally, when an unlock is called, it is created a new triple without a lock, and all the currently locked locks are labelled as unlocked.

Formally, the *initial abstract state* of a function is changed to $s_{init} = \{\{(\emptyset, \emptyset, \varepsilon)\}\}$. During the analysis of a function `g` with an abstract state $s_{\mathtt{g}}$, when a leaf function `f` is called, the abstract state's transformation is changed as follows: $s_{\mathtt{g}} = \{p' \in 2^{2^\Sigma \times 2^\Sigma \times \Pi} \mid \exists\, p \in s_{\mathtt{g}} : p' = \{(A', B', \pi') \in 2^\Sigma \times 2^\Sigma \times \Pi \mid \exists\,(A, B, \pi) \in p : (\pi = \varepsilon \wedge (A', B', \pi') = (A \cup \{\mathtt{f}\}, B, \pi)) \vee [\pi \neq \varepsilon \wedge [((A, B, \pi) \in \mathit{locked}(p) \wedge (A', B', \pi') = (A, B \cup \{\mathtt{f}\}, \pi)) \vee ((A, B, \pi) \notin \mathit{locked}(p) \wedge (A', B', \pi') = (A, B, \pi))]]\}\}$, where *locked* is a function that returns (for a given program path) a set of the (A, B, π) triples where the lock identified by π is currently locked. Further, when a lock identified by the access path π_i is called, the abstract state changes as follows: $s_{\mathtt{g}} = \{p' \in 2^{2^\Sigma \times 2^\Sigma \times \Pi} \mid \exists\, p \in s_{\mathtt{g}} : p' = \{(A', B', \pi') \in 2^\Sigma \times 2^\Sigma \times \Pi \mid (A', B', \pi') = (\emptyset, \emptyset, \varepsilon) \vee ((A', B', \pi') \in p \wedge \pi' \neq \varepsilon) \vee [(A', B', \varepsilon) \in p \wedge \pi' = \pi_i \wedge \mathit{setLocked}(p, \mathit{locked}(p) \cup \{(A', B', \pi')\})]\}\}$, where *setLocked* is a function that labels triples (for a given program path) as those currently locked by their lock. Furthermore, when an unlock identified by the access path π_i is called, the abstract state changes as follows: $s_{\mathtt{g}} = \{p' \in 2^{2^\Sigma \times 2^\Sigma \times \Pi} \mid \exists\, p \in s_{\mathtt{g}} : p' = \{(A', B', \pi') \in 2^\Sigma \times 2^\Sigma \times \Pi \mid (A', B', \pi') = (\emptyset, \emptyset, \varepsilon) \vee ((A', B', \pi') \in p \wedge \pi' \neq \varepsilon \wedge \pi' \neq \pi_i) \vee [(A', B', \pi') \in p \wedge \pi' = \pi_i \wedge \mathit{setLocked}(p, \mathit{locked}(p) \setminus \{(A', B', \pi')\})]\}\}$. Other definitions (e.g. calling and already analysed *nested* function) will be changed analogically.

The *summary* $\chi_{\mathtt{f}} \in 2^{2^\Sigma} \times 2^\Sigma$ of a function `f` is the same as earlier. Only access paths from abstract states are ignored. I.e. $\chi_{\mathtt{f}} = (\boldsymbol{B}, AB)$, where:

- $\boldsymbol{B} = \{B' \in 2^\Sigma \mid \exists\, p \in s_{\mathtt{f}} : \exists\,(A, B, \pi) \in p : B \neq \emptyset \wedge B' = B\}$, where $s_{\mathtt{f}}$ is the abstract state at the end of an interpretation of `f`.
- $AB = \bigcup\limits_{ab \in AB'} ab$, where $AB' = \{ab \in 2^\Sigma \mid \exists\, p \in s_{\mathtt{f}} : \exists\,(A, B, \pi) \in p : ab = A \cup B\}$.

Example 4.2.1. Consider two base cases (*nested atomic section* and *alternating sequence of locks*) in function `f` and `g` from Listing 4.2. There are two lock objects `lockA` and `lockB` that are used simultaneously. Further assume, that `a`, `b`, `c` are leaf nodes of the *call graph*. After the extension of the distinguishment of multiple locks used, the produced abstract states and summaries are as follows:

- $s_{\mathtt{f}} = \{\{(\{\mathtt{a}\}, \{\mathtt{b}\}, \mathtt{lockB}), (\emptyset, \{\mathtt{a}, \mathtt{b}, \mathtt{c}\}, \mathtt{lockA}), (\emptyset, \emptyset, \varepsilon)\}\}$,
  $\chi_{\mathtt{f}} = (\{\{\mathtt{b}\}, \{\mathtt{a}, \mathtt{b}, \mathtt{c}\}\}, \{\mathtt{a}, \mathtt{b}, \mathtt{c}\})$;

- $s_{\mathtt{f}} = \{\{(\emptyset, \{\mathtt{a}, \mathtt{b}\}, \mathtt{lockA}), (\{\mathtt{a}\}, \{\mathtt{b}, \mathtt{c}\}, \mathtt{lockB}), (\emptyset, \emptyset, \varepsilon)\}\}$,
  $\chi_{\mathtt{f}} = (\{\{\mathtt{a}, \mathtt{b}\}, \{\mathtt{b}, \mathtt{c}\}\}, \{\mathtt{a}, \mathtt{b}, \mathtt{c}\})$.

```
void f()
{
    pthread_mutex_lock(&lockA); // {a, b, c}
    a();
    pthread_mutex_lock(&lockB); // {b}
    b();
    pthread_mutex_unlock(&lockB);
    c();
    pthread_mutex_unlock(&lockA);
}
void g()
{
    pthread_mutex_lock(&lockA); // {a, b}
    a();
    pthread_mutex_lock(&lockB); // {b, c}
    b();
    pthread_mutex_unlock(&lockA);
    c();
    pthread_mutex_unlock(&lockB);
}
```

Listing 4.2: A code snippet used to illustrate *distinguishing multiple locks used* during derivation of *sets of functions called atomically*

4.2.3 Distinguishing Multiple Locks in Phase 2

The pairs Ω of functions that should be called atomically are computed the same way as earlier during the *detection of atomicity violations* in **Phase 2**. The analysis again looks for *pairs of functions that should be called atomically* while this is not the case on some path through the CFG. However, this time, there are stored (in addition to a pair of the most recent function calls) all the most recent pairs of function calls locked under individual locks.

An *abstract state* element is the following: $(\mathtt{x}, \mathtt{y}, \Delta, \Lambda) \in \Sigma \times \Sigma \times 2^{\Sigma \times \Sigma} \times 2^{\Sigma \times \Sigma \times \Pi}$, where $(\mathtt{x}, \mathtt{y})$ and Δ are as before. Λ is a set of *locked pairs* of the most recent function calls with their locks' *access paths*. Thus, the abstract state $s \in \boldsymbol{Q}$ is defined as $2^{\Sigma \times \Sigma \times 2^{\Sigma \times \Sigma} \times 2^{\Sigma \times \Sigma \times \Pi}}$. The analysis works as follows. When a function $\mathtt{f}$ is called, it is created a new pair $(\mathtt{x}', \mathtt{y}')$ of the most recent function calls from the previous pair $(\mathtt{x}, \mathtt{y})$ (i.e. $(\mathtt{x}', \mathtt{y}') = (\mathtt{y}, \mathtt{f})$). This pair is also stored to the locked pairs Λ if there are any locks currently locked. Further, it is checked whether the new pair (or just the last call) violates atomicity, and at the same time, it is not locked by any of the stored locks (i.e. $((\mathtt{x}', \mathtt{y}') \in \Omega \wedge (\mathtt{x}', \mathtt{y}') \notin \Lambda) \vee ((\varepsilon, \mathtt{y}') \in \Omega \wedge (\varepsilon, \mathtt{y}') \notin \Lambda)$). When it holds, it is added to the set Δ of pairs that violate atomicity.

More formally, the *initial abstract state* of a function is defined as $s_{init} = \{\{(\varepsilon, \varepsilon, \emptyset, \emptyset)\}\}$. To formalise the analysis of a function, let $\mathtt{f}$ be a called leaf function. Further, let $s_{\mathtt{g}}$ be the abstract state of a function $\mathtt{g}$ being analysed before the function $\mathtt{f}$ is called. After the call of `f`, the abstract state will be changed as follows: $s_{\mathtt{g}} = \{(\mathtt{x}', \mathtt{y}', \Delta', \Lambda') \in \Sigma \times \Sigma \times 2^{\Sigma\times\Sigma} \times 2^{\Sigma\times\Sigma\times\Pi} \mid \exists\,(\mathtt{x},\mathtt{y},\Delta,\Lambda) \in s_{\mathtt{g}} : (\mathtt{x}',\mathtt{y}') = (\mathtt{y},\mathtt{f}) \wedge \Lambda' = \{(\mathtt{x}'_\pi, \mathtt{y}'_\pi, \pi') \in \Sigma\times\Sigma\times\Pi \mid \exists\,(\mathtt{x}_\pi,\mathtt{y}_\pi,\pi)\in\Lambda : (\mathtt{x}'_\pi,\mathtt{y}'_\pi,\pi') = (\mathtt{y}_\pi,\mathtt{f},\pi)\} \wedge \Delta' = \{(\mathtt{x}'',\mathtt{y}'')\in\Sigma\times\Sigma \mid (\mathtt{x}'',\mathtt{y}'')\in\Delta \vee [((\mathtt{x}'',\mathtt{y}'') = (\mathtt{x}',\mathtt{y}') \vee (\mathtt{x}'',\mathtt{y}'') = (\varepsilon,\mathtt{y}')) \wedge (\mathtt{x}'',\mathtt{y}'')\in\Omega \wedge \nexists\,(\mathtt{x}''_\pi,\mathtt{y}''_\pi,\pi'')\in\Lambda' : (\mathtt{x}''_\pi,\mathtt{y}''_\pi) = (\mathtt{x}'',\mathtt{y}'')]\}\}$. Further, when a lock identified by the access path π_i is called, the abstract state is changed as follows: $s_{\mathtt{g}} = \{(\mathtt{x}', \mathtt{y}', \Delta', \Lambda') \in \Sigma \times \Sigma \times 2^{\Sigma\times\Sigma} \times 2^{\Sigma\times\Sigma\times\Pi} \mid \exists\,(\mathtt{x},\mathtt{y},\Delta,\Lambda) \in s_{\mathtt{g}} : (\mathtt{x}',\mathtt{y}',\Delta',\Lambda') = (\mathtt{x},\mathtt{y},\Delta,\Lambda \cup \{(\varepsilon,\varepsilon,\pi_i)\})\}$. Furthermore, when an unlock identified by the access path π_i is called, the abstract state is changed as follows: $s_{\mathtt{g}} = \{(\mathtt{x}', \mathtt{y}', \Delta', \Lambda') \in \Sigma \times \Sigma \times 2^{\Sigma\times\Sigma} \times 2^{\Sigma\times\Sigma\times\Pi} \mid \exists\,(\mathtt{x},\mathtt{y},\Delta,\Lambda) \in s_{\mathtt{g}} : (\mathtt{x}',\mathtt{y}',\Delta',\Lambda') = (\mathtt{x},\mathtt{y},\Delta,\Lambda \setminus \Sigma\times\Sigma\times\{\pi_i\})\}$.

Example 4.2.2. Consider the function `f` from Listing 4.3. There are two lock objects `lockA` and `lockB` that are used simultaneously. Further assume, that `a`, `b` are leaf nodes of the *call graph*. Then assume, that the result of the first phase of the analysis is that a pair of functions `a`, `b` that should be called atomically, i.e., $\Omega = \{(\mathtt{a},\mathtt{b})\}$. Before the extension of the distinguishment of multiple locks used, the analysis would report an atomicity violation of these functions (line 6). That is because the locks are not distinguished, and the unlock of `lockA` (line 5) would unlock everything. On the other hand, after the extension, there are not reported any atomicity violations because the pair of functions is locked using the second lock — `lockB`. The abstract state $s_{\mathtt{f}}$ of the function `f` before line 7 looks like follows: $s_{\mathtt{f}} = \{(\mathtt{a},\mathtt{b},\emptyset,\{(\mathtt{a},\mathtt{b},\mathtt{lockB})\})\}$.

```
1 void f()
2 {
3     pthread_mutex_lock(&lockA); // {}
4     pthread_mutex_lock(&lockB); // {a, b}
5     pthread_mutex_unlock(&lockA);
6     a(); b();
7     pthread_mutex_unlock(&lockB);
8 }
```

Listing 4.3: A code snippet used to illustrate *distinguishing multiple locks used* during *detection of atomicity violations*

4.3 Analysis's Parametrisation

4.4 Local/Global Atomicity Violations

Chapter 5

Implementation of a New Version of Atomer

[[Některé věci částečně převzít z projektové praxe a z Excelu. Zmínit, že byl aktualizován Infer a že byl udělán nějaký refaktoring.]]

5.1 Phase 1 — Detection of Atomic Sets

5.2 Phase 2 — Detection of Atomicity Violations

5.3 Support for C++ and Java Languages

5.4 Additional Locking Mechanisms

Chapter 6

Experimental Evaluation of the New Version of Atomer

[[Některé věci převzít z Excelu a z projektové praxe.]]

6.1 Testing on Hand-Crafted Examples

6.2 Scalability Benchmark

6.3 Evaluation on Validation Programs Derived from Gluon

6.4 Experiments with Real-Life Programs

6.5 Summary of the Evaluation and Future Work

Chapter 7

Conclusion

This thesis started by describing the principles of *static analysis* and *abstract interpretation*. Further, *Facebook Infer* was described — a concrete static analysis framework that uses abstract interpretation — its features, architecture, and existing analysers implemented in this tool. Next, there were described *contracts for concurrency*. The major part of the thesis then aimed at the description of static analyser *Atomer*[1] — proposed and implemented within the author's bachelor's thesis [15] — implemented as a Facebook Infer's module, and that detects *atomicity violations*. It was described its limitations, and thereafter, it was described the proposal and implementation of its extensions and improvements. Lastly, the experimental evaluation of the new features and improvements was depicted, and there were also described other performed experiments and possible future work.

Atomer works on the level of *sequences of function calls*. It is based on the assumption that sequences of function calls executed *atomically once* should probably be executed *always atomically*, and it naturally works with sequences. In the thesis, to improve *scalability*, the use of sequences was *approximated* by *sets*. Further, two new features were implemented: support for *C++* and *Java* languages; and distinguishing *multiple locks used*.

The introduced enhancements were successfully tested on *hand-crafted* programs. It turned out that such innovations improved the *accuracy* and *scalability*. Moreover, Atomer was experimentally evaluated on additional software. Notably, it was evaluated on *real-life Java programs* — *Apache Cassandra* and *Tomcat*. Already fixed and reported *real bugs* were successfully rediscovered. Nevertheless, so far, quite some *false alarms* are reported.

Several other improvements were proposed to reduce the number of false alarms, namely, *parametrisation* of the analysis, support for *interprocedural locks*, or combinations with a *dynamic analysis*. Their implementation and evaluation is currently the work in progress.

Atomer's *accuracy* can be further increased. Some of its limitations and possible solutions are discussed in this thesis, e.g., considering *formal parameters* and distinguishing the *context* of called functions, *ranking* of atomic functions, or focusing on *library containers concurrency restrictions* related to method calls. Further, it is needed to perform more experiments on *real-life* programs to find and report *new bugs*.

[1] The implementation of **Atomer** is available on GitHub as an *open-source* repository (in a branch `atomicity-sets`): `https://github.com/harmim/infer`.

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Appendix A

Contents of the Attached Memory Media

[[Strukturu převzít z bakalářky.]]

Appendix B

Installation and User Manual

[[Převzít z bakalářky a aktualizovat (aktuální informace jsou na Wiki na GitHubu), hlavně přidat použití parametrů analyzátoru. Přidat taky možná instalaci přes Docker.]]

Installation Manual

User Manual