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Scalable Static Analysis Using Facebook Infer

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Abstract
Static analysis has nowadays become one of the most popular ways of catching bugs early in the modern software. However, reasonably precise static analyses do still often have problems with scaling to larger codebases. And efficient static analysers, such as Coverity or Code Sonar, are often proprietary and difficult to openly evaluate or extend. *Facebook Infer* offers a static analysis framework that is open source, extendable, and promoting efficient modular and incremental analysis. In this work, we propose three *inter-procedural* analysers extending the capabilities of Facebook Infer: *Looper* (a resource bounds analyser), *L2D2* (a low-level deadlock detector), and *Atomer* (an atomicity violation analyser). We evaluated our analysers on both smaller hand-crafted examples as well as publicly available benchmarks derived from real-life low-level programs and obtained encouraging results. In particular, L2D2 attained 100 % detection rate and 11 % false positive rate on an extensive benchmark of hundreds of functions and millions of lines of code.

Keywords: Facebook Infer — Static Analysis — Abstract Interpretation — Atomicity Violation — Concurrent Programs — Performance — Worst-Case Cost — Deadlock

Supplementary Material: Looper Repository*** — L2D2 Repository** — Atomer Repository*

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1. Introduction

Bugs are an inherent part of software ever since the inception of the programming discipline. They tend to hide in unexpected places, and when they are triggered, they can cause significant damage. In order to catch bugs early in the development process, extensive automated testing and dynamic analysis tools such as profilers are often used. But while these solutions are sufficient in many cases, they can sometimes still miss too many errors. An alternative solution is a *static analysis*, which has its own shortcomings as well. Like, for example, a high rate of *false positives* and, in particular, quite a big problem with *scalability*.

Recently, Facebook has proposed its own solution for efficient bug finding and program verification called *Facebook Infer* — a highly *scalable compositional* and *incremental* framework for creating *inter-procedural* analyses. Facebook Infer is still under development, but it is in everyday use in Facebook (and several other companies, such as Spotify, Uber, Mozilla, and others) and it already provides many checkers for various kinds of bugs, e.g., for verification of buffer overflow, thread safety, or resource leakage. However, equally importantly, it provides a suitable framework for creating new analyses quickly.

However, the current version of Infer still misses better support, e.g., for *concurrency* or *performance-based* bugs. While it provides a fairly advanced *data race* and *deadlock* analysers, they are limited to Java programs only and fail for C programs, which require more thorough manipulation with locks. Moreover, the only performance-based analyser aims to *worst-case execution time* analysis only, which does not provide a wise understanding of the programs performance.

In particular, we propose to extend Facebook Infer with three analysers: *Looper*, a resource bounds analyser; *L2D2*, a lightweight deadlock checker; and *Atomer*, an atomicity violation checker working on the level of sequence of function calls. In experimental evaluation, we show encouraging results, when even our immature implementation could detect concurrency property violations and infer precise bounds for selected benchmarks, including rather large benchmarks based on real-life code. The development of these checkers has been discussed several times with the developers of Facebook Infer, and it is an integral part of the H2020 ECSEL project Aquas.

2. Facebook Infer

Facebook Infer is an open-source static analysis framework, implemented in OCaml, which is able to discover various types of bugs of the given program, in a *scalable* manner. It is a general *abstract interpretation* [1] framework focused primarily on finding bugs rather than formal verification that can be used to quickly develop new kinds of *compositional* and *incremental* analyses based on the notion of function *summaries*. In theory, a summary is a representation of function's preconditions and postconditions or effects. In practice, it is a custom data structure that allows users to store arbitrary information resulting from function's analysis. Infer does (usually) not compute the summaries during a run of the analysis along the *Control Flow Graph* (CFG) as is done in classical analysers based on the ideas from [2] and [3]. Instead, it analyses a program *function-by-function along the call tree*, starting from its leafs. Hence, a summary of a function is typically analysed without knowing its call context. The summary of the function is then used at all of its call sites. Furthermore, thanks to its incrementality, Infer can analyse individual code changes instead of the whole project, which is more suitable for large and quickly changing codebases where the conventional batch analysis is unfeasible. Intuitively, the incrementality is based on re-using summaries of functions for which there is no change in them nor in the functions (transitively) called from them.

Infer uses a scheduler which determines the order of analysis of individual functions based on a *call graph*. It also checks if it is possible to analyse some functions concurrently, which allows Infer to run in a heavily parallelized manner. In more detail, a call graph is a *directed graph* describing call dependencies between functions. An example of a call graph is shown in Figure 1. Using this figure, we can illustrate the order of analysis in Infer and its incrementality. The underlying analyser starts with the leaf functions P5 and P6 and then proceeds towards the root $\mathtt{P}_{\mathtt{MAIN}}$ while respecting the dependencies represented by the edges. Each subsequent code change then triggers a re-analysis of the directly affected functions only as well as a re-analysis of all the functions up the call chain. For example, if we modify the function P3, Infer will re-analyse only P3, P1, and $\mathtt{P}_{\mathtt{MAIN}}$.

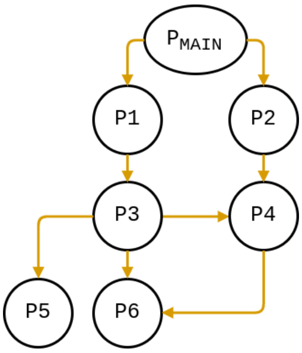


Figure 1. A call graph

Infer supports analysis of programs written in multiple languages including C, C++, Objective-C, and Java and provides a wide range of analyses, each focusing on different types of bugs, such as *Inferbo* (buffer overruns), *RacerD* [4] (data races), or *Starvation* (concurrency starvation and selected types of deadlocks).

3. Worst-Case Cost Analyser

Recently, performance issues have become considerably more widespread in code, leading to a poor user experience. Facebook Infer currently provides the *cost* checker [5] only, which implements a *worst-case execution time* analysis (*WCET*). However, this analysis provides a numerical bound on the time required for the execution of a program only, which can be hard to interpret, and, above all, it is quite imprecise for more complex algorithms, e.g., requiring amortized reasoning. Loopus [6] is a powerful resource bounds analyser, which, to the best of our knowledge, is the only one that can handle *amortized complexity analysis* for a broad range of programs. However, it is limited to intra-procedural analysis only, and the tool itself without an incremental framework is not suitable for large and quickly changing codebases. Hence, we implemented *Looper* – analyser that recasts the powerful analysis of Loopus within Infer which enables the possibility for a more efficient resource bounds analysis.

Bounds inferred by Loopus refer to the number of possible *back jumps* to loop headers, which is an useful metric related to *asymptotic time complexity* as it corresponds to the possible number of executions of instructions inside a loop. The main algorithm relies on an abstract program model called a *difference constraint program* (DCP), an example of which can be seen in Figure 2b.

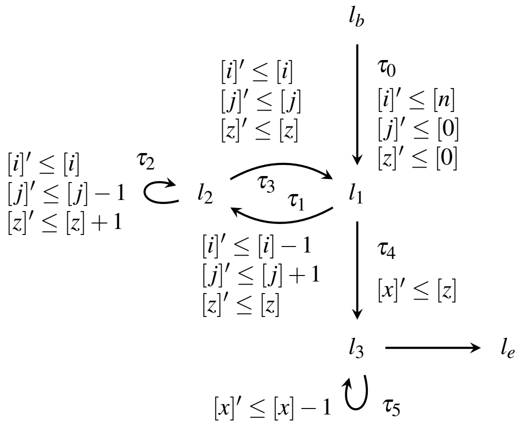
Listing 1. A snippet requiring amortized complexity analysis. The DCP abstraction is shown in Figure 2b. '*' denotes non-determinism. Total cost: $3n$

```
void foo(int n):
    int i = n, j = 0, z = 0;
l1: while (i > 0):
        i--; j++;
l2:     while (j > 0 && *) j--; z++;
    int x = z;
l3: while (x > 0) x--;
```

Each transition τ of a DCP has a *local bound* τ_v,

Call	Evaluation and Simplification
$\mathcal{TB}(\tau_5)$	$\rightarrow \mathtt{Incr}([x])+$ $\mathcal{TB}(\tau_4) \times \max(\mathcal{VB}([z])+0,0)$ $\rightarrow 0+1\times\max([n]+0,0)=[n]$
$\mathcal{VB}([z])$	$\rightarrow \mathtt{Incr}([z])+\max(\mathcal{VB}(0)+0)=[n]$
$\mathtt{Incr}([z])$	$\rightarrow \mathcal{TB}(\tau_2)\times 1=[n]$
$\mathcal{TB}(\tau_2)$	$\rightarrow \mathtt{Incr}([j])+\mathcal{TB}(\tau_0)\times 0$ $\rightarrow [n]+1\times 0=[n]$
$\mathtt{Incr}([j])$	$\rightarrow \mathcal{TB}(\tau_1)\times 1=[n]$
$\mathcal{TB}(\tau_1)$	$\rightarrow \mathtt{Incr}([i])+\mathcal{TB}(\tau_0)\times\max([n]+0,0)$ $\rightarrow 0+1\times[n]=[n]$

(a) A simplified computation of the bound for τ_5. $\mathtt{Incr}([x])$ and $\mathtt{Incr}([i])$ are 0 as there are no transitions that increase the value of $[x]$ or $[i]$. $\mathcal{TB}(\tau_0)$ and $\mathcal{TB}(\tau_4)$ are 1 as they are not part of any loop.



(b) An abstraction obtained from Listing 1. Each transition is denoted by a set of invariant inequalities.

Figure 2

i.e., a variable v that *locally* limits the number of executions of the transition τ. For example, the variable j in Figure 2b limits the number of consecutive executions of the transition τ_2.

The bound algorithm is based on the idea of reasoning about *how often* and *by how much* might the local bound of a transition τ increase, which affects the number of executions of τ. The computation interleaves calls to two procedures:

1. $\mathcal{VB}$ – computes a *variable bound* expression in terms of program parameters which bounds the value of the variable v.
2. $\mathcal{TB}$ – computes a bound on the number of times that a transition τ can be executed. Transitions that are not part of any loop have the transition bound 1.

The $\mathcal{TB}$ procedure is defined in the following way:

$$\mathcal{TB}(\tau) = \mathtt{Incr}(\tau_v) + \mathtt{Resets}(\tau_v)$$

The $\mathtt{Incr}(\tau_v)$ procedure represents *how often* and *by how much* might the local bound τ_v increase:

$$\mathtt{Incr}(\tau_v) = \sum_{(\mathtt{t},\mathtt{c})\in\mathcal{I}(\tau_v)} \mathcal{TB}(\mathtt{t}) \times \mathtt{c}$$

$\mathcal{I}(\tau_v)$ is the set of transitions $\mathtt{t}$ that increase the value of τ_v by $\mathtt{c}$. $\mathtt{Resets}(\tau_v)$ represent the possible resets of the local bound τ_v to some arbitrary values which also add to the total amount by which τ_v and consequently $\mathcal{TB}(\tau)$ might increase:

$$\mathtt{Resets}(\tau_v) = \sum_{(\mathtt{t},\mathtt{a},\mathtt{c})\in\mathcal{R}(\tau_v)} \mathcal{TB}(\mathtt{t}) \times \max(\mathcal{VB}(\mathtt{a})+\mathtt{c},0)$$

Above, $\mathcal{R}(\tau_v)$ is the set of transitions $\mathtt{t}$ that reset the value of the local bound τ_v to $\mathtt{a}+\mathtt{c}$ where $\mathtt{a}$ is a variable.

The remaining $\mathcal{VB}(\mathtt{v})$ procedure is defined as:

$$\mathcal{VB}(\mathtt{v}) = \mathtt{Incr}(\mathtt{v}) + \max_{(\mathtt{t},\mathtt{a},\mathtt{c})\in\mathcal{R}(\mathtt{v})} (\mathcal{VB}(\mathtt{a})+\mathtt{c})$$

It picks the maximal value of all possible resets of $\mathtt{v}$ as the initial value and increases it by the value of $\mathtt{Incr}(\mathtt{v})$. Note that the procedure returns $\mathtt{v}$ itself if it is a program parameter or a numeric constant.

The complete bound algorithm is then the mutual recursion of the procedures $\mathcal{TB}$ and $\mathcal{VB}$. The main reason why Loopus scales so well with this approach is *local* reasoning: it does not rely on any global program analysis and is able to obtain complex invariants such as $x \leq \max(m1, m2) + 2n$. These invariants are not expressible in common abstract domains such as *octagon* or *polyhedra*, which would lead to a less precise result. This approach is also *demand-driven* (Figure 2a), i.e. it only performs necessary recursive calls and does not compute all possible invariants. For a full *flow* and *path sensitive* algorithm and its extension refer to [6].

Figure 2a presents an example computation of the transition bound of τ_5 from the DCP in Figure 2b, which corresponds to Listing 1. This code demonstrates the need for amortized complexity analysis as the worst-case cost of the l_2 loop can indeed be n. However, its amortized cost is 1 as the total number of iterations of l_2 (total cost) is also equal to n due to the local bound j, which is bounded by n. Loopus is able to obtain the bound of n instead of n^2 for the inner loop l_2 unlike many other tools. Another challenge is the computation of the bound for the loop l_3. It is easy to infer z as the bound, but the real challenge lies in expressing the bound in terms of program parameters. Thus, the real task is to obtain an invariant of the form $z \leq \mathtt{expr}(n)$ where $\mathtt{expr}(n)$ denotes an expres-

Table 1. An experimental evaluation of *Looper*. Benchmarks are publicly available[1].

	Bound	Inferred bound		Time [s]	
		Looper	Cost	Looper	Cost
#1	n	$2n$	n^2	0.3	0.4
#2	$2n$	$2n$	$5n$	0.5	0.4
#3	$4n$	$5n$	∞	0.8	1.4
#4	$*n^2$	n^2	∞	0.6	0.9
#5	$2n$	$2n$	$12n$	0.3	0.5
#6	$*n$	n	∞	0.6	0.7
#7	$2n$	$2n$	∞	0.4	1
#8	$2n$	$2n$	∞	0.7	1.8

sion over program parameters, n in this case. Loopus is able to obtain the invariant $z \leq n$ simply with the $\mathcal{VB}$ procedure and to infer the bound n for the loop l_3.

The implementation of $\mathcal{TB}$ and $\mathcal{VB}$ is quite straightforward in a functional paradigm (OCaml). We first convert the native CFG used by Infer into a DCP used by Loopus' abstraction. In particular, we leverage the AI framework and symbolically execute the program yielding a transition system. Further, we had to implement the abstraction algorithm and an algorithm which computes local bounds. We further extended the basic algorithm with several extensions which improve its precision such as a reasoning based on so called *reset chains* or an algorithm that converts the standard DCP into a *flow-sensitive* one by variable renaming. For more details about these extensions, refer to [6]. The current implementation is still limited to intra-procedural analysis as the original Loopus. However, we already have a conceptual idea based on substitution of the formal parameters in a symbolic bound expression stored in a summary with the variable bounds of arguments at a callsite resulting in, albeit less precise, but scalable solution. We should also be able to obtain the symbolic return value through the $\mathcal{VB}$ procedure and then use it at a call site in a similar way. We are aware that this reasoning is limited to functions without pointer manipulation but it should be a step in the right direction.

Table 1 presents experimental results of our current implementation on selected examples from the dissertation [6]. We compared the results of *Looper* (Loopus in Infer) with the *Cost* analyser mentioned in the introduction of this section. For *Cost* we have simplified the reported bounds to the worst-case asymptotic complexity instead of the cost.

4. Deadlock Analyser

According to [7], deadlock is perhaps the most common concurrency error that might occur in almost all parallel programming paradigms including both shared-memory and distributed memory. Detecting deadlocks during testing is very hard due to many possible interleavings among threads. Of course, one can use extrapolating dynamic analysers and/or techniques such as noise injection or systematic testing [8] to increase chances of finding deadlocks, but such techniques decrease the scalability of the testing process and can still have problems discovering some errors. That is the reason why many static detectors were created, but most of them are quite heavy-weight and do not scale well. However, there are few that meet the scalability condition, like the *starvation* analyser implemented in Facebook Infer. But, the problem of this analyser is that it uses a heuristic based on using the class of the root of the *access path*[2] of a lock, and so it does not handle pure C locks. Another, that is worth mentioning, is the RacerX analyser [9], which is based on counting so-called *locksets*, i.e., sets of locks currently held. RacerX uses interprocedural, flow-sensitive, and context-sensitive analysis. This means that each function needs to be reanalysed in a new context, which reduces the scalability. Hence, we have decided to develop a new context-insensitive analysis (only very loosely inspired by RacerX), which will be faster and more scalable. We have implemented this analysis in our *Low-Level Deadlock Detector (L2D2)*, the principle of which will be illustrated by the example in Listing 2 (a full description of the algorithm with all its optimisations is beyond the scope of this paper).

L2D2 works in two phases. In the first phase, it computes a summary for each function by looking for lock and unlock events present in the function. An example of a lock and an unlock event is illustrated in Listing 2 at lines 22 and 27. If a call of a user-defined function appears in the analysed code during the analysis, like at line 26 of our example, the analyser is provided with a summary of the function if available. Otherwise, the function is analysed on demand (which effectively leads to analysing the code along the call tree, starting at its leaves, as usual in Facebook Infer). The summary is then applied to an abstract state at the call site. Hence, in our example, the summary of `foo` will be applied to the abstract state of `thread1`. More details on what the summaries look like and how they are computed will be given in Section 4.1.

[1] https://bit.ly/2uORslv

[2] Infer uses access paths for naming heap locations via the paths used to access them, e.g., `x.f.g` (`x` is the root).

Listing 2. A simple example capturing a deadlock between two global locks in the C language using the POSIX threads execution model

```
void foo() {
    pthread_mutex_lock(&lock2);
void *thread1(...) {
    pthread_mutex_lock(&lock1);
      ⋮
    foo();
    pthread_mutex_unlock(&lock1);
void *thread2(...) {
    pthread_mutex_lock(&lock2);
      ⋮
    pthread_mutex_lock(&lock1);
```

In the second phase, L2D2 looks through all the computed summaries of the analysed program and concentrates on so-called *dependencies* that are part of the summaries. The dependencies record that some lock got locked at a moment when another lock was still locked. L2D2 interprets the obtained set of dependencies as a relation, computes its transitive closure, and reports a deadlock if some lock depends on itself in the transitive closure.

If we run L2D2 on our example, it will report a possible deadlock due to the cyclic dependency between `lock1` and `lock2` that arises if thread 1 holds `lock1` and waits on `lock2` and thread 2 holds `lock2` and waits on `lock1`. This is caused by dependencies `lock1→lock2` and `lock2→lock1` in the summaries of `thread1` and `thread2` (see Listing 3).

4.1 Computing Function Summaries

We now describe the structure of the summaries used and the process of computing them. To detect potential deadlocks, we need to record information that will allow us to answer the following questions:

(1) What is the state of the locks used in the analysed program at a given point in the code?
(2) Could a cyclic dependency on pending lock requests occur?

To answer question (1), we compute sets *lockset* and *unlockset*, which contain the currently locked and the currently unlocked locks, respectively. These sets are a part of the *postconditions* of functions and record what locks are locked/unlocked upon returning from a function, respectively. Further, we also compute sets *locked* and *unlocked* that serve as a *precondition* for a given function and contain locks that should be locked/unlocked before calling this function. When analysing a function, the sets are manipulated as shown in Figure 4.

Listing 3. Summaries of the functions in Listing 2

```
foo()
  PRECONDITION: { unlocked={lock2} }
  POSTCONDITION: { lockset={lock2} }
thread1(...)
  PRECONDITION: { unlocked={lock1, lock2} }
  POSTCONDITION: { lockset={lock2},
    dependencies={lock1->lock2} }
thread2(...)
  PRECONDITION: { unlocked={lock1, lock2} }
  POSTCONDITION: { lockset={lock1, lock2},
    dependencies={lock2->lock1} }
```

Listing 4. Rules for summary computation

```
lockset:
  lock(l) → lockset := lockset ∪ {l}
  unlock(l) → lockset := lockset - {l}
unlockset:
  lock(l) → unlockset := unlockset - {l}
  unlock(l) → unlockset := unlockset ∪ {l}
locked:
  if(lock(l) is the first operation in f)
    unlocked_f := unlocked_f ∪ {l}
unlocked:
  if(unlock(l) is the first operation in f)
    locked_f := locked_f ∪ {l}
```

Each summary contains also a set of *dependencies* using which we can answer question (2). Extraction of the dependencies is called upon every lock acquisition and iterates over every lock in the current lockset, emitting the ordering constraint produced by the current acquisition. For example, if `lock2` is in the current lockset and `lock1` has just been acquired, the dependency `lock2→lock1` will be emitted, as we can see in Listing 2 in the function `thread2`.

The above described basic computation of the dependencies would, however, be very imprecise and lead to many false alarms. The imprecision is caused by invalid locksets. The main reasons for imprecision of the locksets are imprecision in dealing with conditionals (all outcomes are considered as possible), with function calls (missing context), and with lock aliasing (any aliasing is considered to be possible).

Next, as we mentioned in the beginning of this section, if a function call appears in the analysed code, we have to apply a summary of the function to the abstract state at the call site. Given a callee g, its lockset L_g, unlockset U_g, and a caller f, its lockset L_f, unlockset U_f, and dependencies D_f, we:

(1) Update the summary of g by replacing formal parameters with actual ones in case that locks were passed to g as parameters. In the example below, you can notice that `lock4` will be replaced by `lock2` in the summary of g.
(2) Update the precondition of f:
if$(\exists l : l \in \mathit{unlocked}_g \wedge l \notin \mathit{unlockset}_f)$

```
        add lock $l$ to $unlocked_f$
if($\exists l : l \in locked_g \wedge l \notin lockset_f$)
        add lock $l$ to $locked_f$
```

(3) Update D_f by adding new dependencies for all locks in L_f with locks which were locked in g.

However, a problem occurs if some of the locks which were acquired in g were also released there. This is illustrated in the example below.

```
void f():
    pthread_mutex_lock(&lock2);
    g(&lock2);
void g(pthread_mutex_t *lock4):
    pthread_mutex_lock(&lock3);
    pthread_mutex_unlock(lock4);
    pthread_mutex_lock(&lock1);
        ⋮
    pthread_mutex_unlock(&lock1);
    pthread_mutex_unlock(&lock3);
```

In that case, L_g will not contain these locks, and we have no information about them. To cope with problem, we have yet another set in the summaries whose semantics is similar to the semantics of the lockset except that the unlock statement does not remove locks from it. In our example, this set would contain `lock3` and `lock1`. Moreover, there is still one problem left. What if the lock from the current lockset was unlocked in the callee before we locked another lock there? Then we would emit the wrong dependency `lock2→lock1`. In order to avoid this problem, we create `unlock→lock` type dependencies in the summaries, that can be used to safely determine the order of operations in the callee. This finally ensures that the only newly created correct dependency in our example will be `lock2→lock3`.

(4) Update L_f: $L_f = (L_f \setminus U_g) \cup L_g$

(5) Update U_f: $U_f = (U_f \setminus L_g) \cup U_g$

4.2 Experimental Evaluation

We performed experiments using a benchmark of 1002 concurrent C programs derived from the Debian GNU Linux distribution.The entire benchmark is available online at GitLab[3]. These programs were originally used for an experimental evaluation of Daniel Kroening's static deadlock analyser [10] implemented in the CPROVER framework.

This benchmark set consists of 11.4 MLOC. Of all the programs, 994 are deadlock-free and 8 of them contain a deadlock. Our experiments were run on a CORE i7-7700HQ processor at 2.80 GHz running Ubuntu 18.04 with 64-bit binaries. The CPROVER experiments were run on a Xeon X5667 at 3 GHz running Fedora 20 with 64-bit binaries. In case of CPROVER, the memory and the CPU time were restricted to 24 GB and 1800 seconds per benchmark, respectively.

Table 2. Results for programs without a deadlock (t/o — timed out, m/o — out of memory)

	proved	alarms	t/o	m/o	errors
CPROVER	**292**	114	453	135	0
L2D2	**810**	104	0	0	80

Both our analyser and CPROVER correctly report all 8 potential deadlocks in the benchmarks with known issues. A comparison of results for deadlock-free programs can be seen in Table 2.

As one can see, L2D2 reported false alarms for 104 deadlock-free benchmarks which is by 10 less than CPROVER. A much larger difference can be seen in cases where it was proved that there was no deadlock. The difference here is 518 examples in favor of our analyser. In case of L2D2, we have 80 compilation errors that were caused by syntax that Infer does not support. The biggest difference between our analyser and CPROVER is the runtime. While our analyser needed approximately 2 hours to perform the experiments, CPROVER needed about 300 hours.

There is still space for improving our analysis by reducing the number of alarms, which are mainly caused by false dependencies as mentioned in Subsection 4.1 (4^{th} paragraph). Hence, to eliminate false positives, we need some techniques to eliminate false dependencies. In our implementation of L2D2, we use a number of heuristics that try to reduce the imprecision. An example is that if a locking error occurs (double lock acquisition), then L2D2 sets the current lockset to empty, and adds the currently acquired lock to the lockset (we can safely tell that this lock is locked), thereby eliminating any dependencies that could result from the locking error. More precise description of these heuristics is beyond the scope of the paper.

5. Atomicity Violations Analyser

In *concurrent programs* there are often *atomicity requirements* for execution of specific sequences of instructions. Violating these requirements may cause many kinds of problems, such as an unexpected behaviour, exceptions, segmentation faults, or other failures. *Atomicity violations* are usually not verified by compilers, unlike syntactic or some sorts of semantic rules. Atomicity requirements, in most cases, are not even documented. It means that typically only programmers must take care of following these requirements. In general, it is very difficult to avoid errors in *atomicity-dependent programs*, especially in large projects, and even harder and time-consuming is finding and fixing these errors.

[3] https://bit.ly/2WJBQLQ

Listing 5. An example of a contract violation

```
void replace(int *array, int a, int b):
    int i = index_of(array, a);
    if (i >= 0) set(array, i, b);
```

In this section we propose an implementation of a *static analyser for finding atomicity violations*. In particular, we concentrate on an *atomic execution of sequences of function calls*, which is often required, e.g., when using certain library calls.

5.1 Contracts for Concurrency

The proposal of a solution is based on the concept of *contracts for concurrency* described in [11]. These contracts allow one to define *sequences of functions* that are required to be *executed atomically*. The proposed analyser itself (**Atomer**) is able to automatically derive candidates for such contracts, and then to verify whether the contracts are fulfilled.

In [11], 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 the set $\mathbb{R}$ of *clauses* where each clause $\varrho \in \mathbb{R}$ is a regular expression over $\Sigma_{\mathbb{M}}$. A *contract violation* occurs if any of the sequences represented by the contract clauses is interleaved with an execution of functions from $\Sigma_{\mathbb{M}}$.

Consider an implementation of a function that replaces item a in an array by item b, illustrated in Listing 5. The contract for this specific scenario contains clause ϱ_1, which is defined as follows:

$$(\varrho_1)\ \texttt{index\_of set}$$

Clause ϱ_1 specifies that every execution of index_of followed by an execution of set should be atomic. The index of an item in an array is acquired, and then the index is used to modify the array. Without atomicity, a concurrent modification of the array may change a position of the item. The acquired index may then be invalid when set is executed.

In [11] there is described a proposal and an implementation of a *static validation for finding atomicity violations*, which is based on *grammars* and *parsing trees*. The authors of [11] implemented the stand-alone prototype tool[4] for analysing programs written in Java, which led to some promising experimental results but the *scalability* of the tool was still limited. Moreover, the tool from [11] is no more developed. That is why we decided to get inspired by [11] and reimplement the analysis in *Facebook Infer* redesigning it in accordance with the principles of Infer, which should make it more *scalable*. In the end, due to adapting the analysis for the context of Infer, our implementation is significantly different from [11] as presented in Sections 5.2 and 5.3. Furthermore, unlike [11], the implementation aims at programs written in C/C++ languages using *POSIX Threads (Pthreads)* locks for *synchronisation of concurrent threads*.

In Facebook Infer there is already implemented an analysis called Lock Consistency Violation, which is part of *RacerD* [4]. The analysis finds atomicity violations for writes/reads on single variables that are required to be executed atomically. Atomer is different, it finds *atomicity violations for sequences of functions* that are required to be executed atomically, i.e., it checks whether contracts for concurrency hold.

The proposed solution is divided into two parts (*phases of the analysis*):

Phase 1 Detection of *atomic sequences*, which is described in Section 5.2.

Phase 2 Detection of *atomicity violations*, which is described in Section 5.3.

5.2 Detection of Atomic Sequences

Before the detection of *atomicity violations* may begin, it is required to have *contracts* introduced in Section 5.1. **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 that are 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.

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)** of sets of function calls such that: **A** is a *reduced sequence* of function calls that appear between the beginning of the function being analysed and the first lock or between an unlock and a subsequent lock (or the end of the function being analysed), and **B** is a *reduced sequence* of function calls that follow the calls from **A** and that appear between a lock and an unlock (or the end of the function being analysed). Here, by a *reduced sequence* we mean a sequence in which the first appearance of each function is recorded only. The reason is to ensure *finiteness* of the sequences and of the analysis. The *summary* then consists of (i) the set of all the **B** sequences and (ii) the set of *concatenations* of all the **A** and **B** sequences with removal of duplicate function calls. The latter is recorded for the purpose of analysing functions higher in the call hierarchy since locks/unlocks can appear in such a higher-level function.

[4] https://github.com/trxsys/gluon

Listing 6. An example of a code for an illustration of the derivation of sequences of functions called atomically

```
void g(void):
    f1(); f1();
    pthread_mutex_lock(&lock);
    f1(); f1(); f2();
    pthread_mutex_unlock(&lock);
    f1(); f1();
    pthread_mutex_lock(&lock);
    f1(); f3();
    pthread_mutex_unlock(&lock);
    f1();
    pthread_mutex_lock(&lock);
    f1(); f3(); f3();
    pthread_mutex_unlock(&lock);
```

For instance, analysis of the function g from Listing 6 (assuming *Pthreads* locks and existence of the initialized global variable lock of the type pthread_mutex_t) produces the following sequences:

$$\overbrace{\texttt{f1}\ \cancel{\texttt{f1}}}^{\mathbf{A_1}}\overbrace{(\texttt{f1}\ \cancel{\texttt{f1}}\ \texttt{f2})}^{\mathbf{B_1}}\,|\,\overbrace{\texttt{f1}\ \cancel{\texttt{f1}}}^{\mathbf{A_2}}\overbrace{(\texttt{f1}\ \texttt{f3})}^{\mathbf{B_2}}\,|\,\overbrace{\cancel{\texttt{f1}}}^{\cancel{\mathbf{A_3}}}\ \overbrace{(\cancel{\texttt{f1}\ \texttt{f3}\ \texttt{f3}})}^{\cancel{\mathbf{B_3}}}$$

The parentheses are used to indicate an atomic sequence. The strikethrough of the functions f1 and f3 denotes a removal of already recorded function calls in the **A** and **B** sequences. The strikethrough of the entire sequence f1 (f1 f3 f3) means discarding sequences already seen before. The derivated sets for the function g are then as follows: (i) {(f1 f2), (f1 f3)}, i.e., $\mathbf{B_1}$ and $\mathbf{B_2}$, (ii) {f1 f2 f3}, i.e., concatenation of $\mathbf{A_1}$, $\mathbf{B_1}$, $\mathbf{A_2}$, and $\mathbf{B_2}$ with removal of duplicate function calls.

Further, we show how the function h from Listing 7 would be analysed using the result of the analysis of the function g. The result of the analysis of the nested function is used as follows. When calling an already analysed function, one plugs all the sequences from the second component of its summary into the current **A** or **B** sequence. So the analysis of the function h produces the following sequence: f1 g ~~f1~~ f2 f3 (g f1 f2 f3). The derivated sets for the function h are as follows: (i) {(g f1 f2 f3)}, (ii) {f1 g f2 f3}.

The above detection of atomic sequences has been implemented and successfully verified on a set of sample programs created for this purpose. The derived sequences of calls assumed to execute atomically, i.e., the **B** sequences, from the summaries of all analysed functions are stored into a file, which is used during **Phase 2**, described below. There are some possibilities for further extending and improving **Phase 1**, e.g., working with nested locks, distinguishing the different locks used (currently, we do not distinguish between the locks at all), or extending the detection for other types of locks for synchronisation of concurrent threads/processes. On the other hand, to further enhance the *scalability*, it seems promising to replace working with the **A** and **B** sequence by working with sets of calls: sacrificing some precision but gaining the speed.

Listing 7. An example of a code for an illustration of the derivation of sequences of functions called atomically with nested function call

```
void h(void):
    f1(); g();
    pthread_mutex_lock(&lock);
    g();
    pthread_mutex_unlock(&lock);
```

Listing 8. Atomicity violation

```
void g(void):
    f1();
    pthread_mutex_lock(&lock);
    f2(); f3();
    pthread_mutex_unlock(&lock);
    f4();
void h(void):
    f1(); f2(); f3(); f4();
```

5.3 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 while this is not the case on some path through the CFG.

For instance, assume the functions g and h from Listing 8. The set of atomic sequences of the function g is {(f2 f3)}. In the function h, an atomicity violation is detected because the functions f2 and f3 are not called atomically (under a lock).

Implementation of this phase and its experimental evaluation is currently in progress. Based on the results, we will tune **Phase 1** as well.

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