Automatically finding atomic regions for fixing bugs in Concurrent programs

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Abstract. This paper presents a technique for automatically constructing a fix for buggy concurrent programs: given a concurrent program that does not satisfy user-provided assertions, we infer atomic blocks that fix the program. An atomic block protects a piece of code and ensures that it runs without interruption from other threads. Our technique uses a verification tool as a subroutine to find the smallest atomic regions that remove all bugs in a given program. Keeping the atomic regions small allows for maximum concurrency. We have implemented our approach in a tool called ATOMICINF. A user of ATOMICINF can choose between strong and weak atomicity semantics for the inferred fix. While the former is simpler to find, the latter provides more information about the bugs that got fixed.

We ran ATOMICINF on several benchmarks and came up with the smallest and the most precise atomic regions in all of them. We implemented an earlier technique to our setting and observed that ATOMICINF is 1.7 times faster on an average as compared to an earlier approach.

1 Introduction

An important part of the debugging process is to come up with a repair that fixes the bug under investigation. After a candidate repair is formulated, not only must one reason that the fix removes the bug, but also that it does not introduce new bugs in the program. Thus, evaluating a repair requires understanding of the program as a whole, not just the executions that reveal the current bug. Consequently, any automation in the process of formulating and evaluating a fix would be welcome.

At first, this debugging problem seems to be a good match for verification tools that are prepared to reason over many (or all) program behaviors. However, the process of formulating a fix can be difficult to automate. For instance, any program with an assertion `assert(!error)` can be “fixed” by inserting the statement `error=false` right before the assert. However, such repairs are clearly of no practical use.

In order to get across this challenge of meaningless repairs, we focus on a restricted problem of automated program repair. In particular, we focus on fixing concurrent programs by introducing extra synchronization that restricts the set of interleavings possible in the program. Given a specification, we infer a
fix that removes all of the bad interleavings of the program while minimizing the
set of good interleavings that get removed by the fix. This problem definition has
several advantages. First, a fix is not allowed to introduce new behaviors, i.e., any
execution of the fixed program is also a valid execution of the original program.
This rules out, for instance, the trivial repair mentioned previously. Second, we
remove all bad interleavings, which implies that the resulting program will satisfy
the specification. Third, by minimizing the set of good interleavings removed, we
allow maximum concurrency in the program and avoid significantly degrading
the performance and responsiveness of the program. Fourth, the specification is
supplied by the user, allowing one to target the repair towards certain (user-
defined) properties.

We restrict the space of interleavings by introducing extra synchronization
in the program in the form of atomic blocks. Atomic blocks are a convenient
way of expressing synchronization. Previous work shows that programs that use
atomic blocks are easier to understand than ones that use locks [17]. An atomic
block is used to enclose a piece of code that restricts how that code interacts
with concurrently executing threads. The exact semantics depends on the type of
atomic block used. A strong atomic block ensures that the enclosed code executes
in complete isolation of the rest of the program. A weak atomic block ensures
that the enclosed code executes in isolation of other weak atomic sections. We
allow a user to pick which kind of atomic blocks to use for the fix. A fix using
strong atomic blocks is easier to find, but a fix using weak atomic blocks usually
reveals more information about the bug getting fixed. Furthermore, it is easy to
realize weak atomic blocks using locks [4].

Our approach works as follows. We accept, as input, a program with assert-
tions. We assume that the program specification is fully captured by the asserts.
(Any safety property can be captured using assertions.) Furthermore, we as-
sume that all executions of the program in which threads do not interleave (i.e.,
the threads execute sequentially, one after the other) are correct. This is an im-
portant assumption because otherwise, no set of atomic blocks could repair the
program. Next, we use any off-the-shelf verification tool to iteratively reveal
more and more buggy traces in the program until we converge on a fix. Because
queries to the verification tool can be expensive, we minimize the number of
buggy traces required.

When a user selects strong atomicity, we guarantee that the reported fix is
the smallest in terms of the number of program points protected by the atomic
block. However, the search for the smallest weak atomic blocks turns out to be
too expensive. Thus, when a user selects weak atomicity, we employ a crucial
optimization. We first find the smallest fix $F$ under strong atomicity and then
restrict the search for weak atomic blocks to those which are supersets of $F$.
While this implies that the fix may not be the smallest, we still guarantee that
it is a minimal extension of $F$. Furthermore, our experiments reveal that this
optimization does not compromise the quality of the fix.

The key contributions of this paper are as follows:
We give an efficient approach for finding a smallest fix $F$ under strong atomicity as well as a minimal extension of $F$ under weak atomicity.

Our approach is completely driven by user-supplied properties, as opposed to previous work that relies on symptoms such as data races and atomicity violations to root-cause a bug [5, 18]. Hence, our technique does not get distracted by benign data races.

Our experiments show that we are able to find the best fix on a variety of benchmarks.

The rest of the paper is organized as follows. Section 2 gives an overview of our technique. Section 2.1 describes the related work. Sections 3 to 5 describe our techniques in detail with algorithms and proofs. Section 6 mentions experimental results.

2 Overview

This section gives an overview of AtomicInf. It accepts as input a concurrent program with assertions. We use the term bug to refer to an execution that ends in an assertion violation.

We consider two semantics for atomic blocks. A strong atomic block $\text{satomic}(S)$ guards a region of code $S$ and ensures that it runs in complete isolation with respect to other threads. Intuitively, this means that context switching is disabled while $S$ is executing. A weak atomic block $\text{watomic}(S)$ ensures that $S$ runs in isolation with respect to other weak atomic blocks. One simple (semantics preserving) implementation of weak atomic blocks is to use a single global lock $l$ and replace each block $\text{watomic}(S)$ with $(\text{acquire}(l); S; \text{release}(l))$. In this sense, it is much easier to realize weak atomic blocks in a language runtime and people have proposed efficient implementations for them [4].

Consider the banking program shown in Fig. 1. The procedure transfer transfers a given amount from one account to another account. The procedure seize sets the balance in a given account to zero. The thread thread1 attempts to transfer 200 units from acc1 to acc2; thread2 tries to set the balance in acc1 to zero; and thread3 tries to transfer 100 units from acc2 to acc1. The program is buggy: for example, immediately after thread1 checks whether sufficient amount is available in acc1 in transfer, thread3 starts and runs to completion, then thread2 also runs to completion, setting the balance in acc1 to 0. Next, when thread1 finishes, the amount in acc1 is negative. The appropriate fixes for the program are the following: (1) enclose the body of transfer in a strong atomic block; or (2) enclose the bodies of both transfer and seize in weak atomic blocks. Note that the latter solution is more informative because it points out the conflicting concurrent accesses. AtomicInf can find these fixes automatically.

Let VeriTool refer to any verification tool that gives us the ability of controlling the places where context switches are allowed. Let $P$ be the buggy concurrent program. We feed $P$ to VeriTool and get a trace $t$ that witnesses a buggy execution of $P$. We examine the set of program locations where the trace
struct Account {
    int amount;
} acc1, acc2, acc3;

void seize(Account *acc) {
    acc->amount = 0;
}

void thread1() {
    transfer(&acc1,&acc2,200);
}

void thread2() {
    seize(&acc1);
}

void thread3() {
    transfer(&acc2,&acc1,100);
}

int transfer(Account* src, Account* dst, int amount) {
    if(src->amount >= amount) {
        int temp = src->amount;
        temp = temp - amount;
        src->amount = temp;
        temp = dst->amount;
        temp = temp + amount;
        dst->amount = temp;
        return 1;
    }
    return 0;
}

void main() {
    acc1.amount = acc2.amount = acc3.amount = 200;
    t1 = async thread1();
    t2 = async thread2();
    t3 = async thread3();
    join(t1); join(t2); join(t3);
    assert(acc1.amount == 0 || acc1.amount == 100);
}
int transfer(Account src, Account dst, int amount)
{
    nonatomic {
        if(src.amount>=amount)
        {
            int temp = src.amount;
            temp = temp - amount;
            src.amount = temp;
            temp = dst.amount;
            temp = temp + amount;
            dst.amount = temp
            return 1;
        }
    }
    return 0;
}

void thread2()
{
    int temp=0;
    nonatomic{
        temp = acc1.account;
        temp += acc2.account;
        assert(temp == 400);
    }
}

For weak atomicity, we follow a similar iterative process, but mine the error traces for more information. In particular, we look for pairs of locations \((l_1, l_2)\) such that a buggy trace takes a context switch at \(l_1\) and passes through \(l_2\). We can rule out this error trace by placing both locations inside a weak atomic block. To reduce the search space, we first find a solution under strong atomicity and then use it as a starting point for finding a solution under weak atomicity. In many cases, these solutions are very similar, thus, reducing the number of iterations required for weak atomicity. For the banking example, once the strong atomicity fix is found, extending it to a weak atomicity fix only requires enclosing the seize procedure in a weak atomic block.

We now illustrate the property-guided nature of AtomicInf. For the program of Fig. 4, the fix reported by AtomicInf allowed a context switch at line 10 as shown in Fig. 2(a). On closer inspection, this is a valid solution; it says that operations of debiting amount from src, and crediting to dst need to be individually atomic, but it is fine for other operations to execute between them.

As a further test, we changed the implementation of the second thread to what is shown in Fig. 2(b). It checks that the corpus of money in the two accounts remains constant. Because this is done in a thread, the assertion can fire any time during the program’s execution. In this case, AtomicInf proposes that the entire body of transfer needs to be inside a single strong atomic block; it is no longer safe to interleave operations between the debit and credit of transfer.

2.1 Related Work

Automatic repair of programs has been studied earlier, both for sequential programs \([2, 6, 9, 13]\) as well as for concurrent programs \([3, 8, 15, 21, 22]\). Previous
work on sequential programs has focused on formulating program repair as a
two-player game, where one of the players tries to make sure that the program
doesn’t fail. A winning strategy for this player is the repair. One limits the vo-
cabulary of the player (to, for instance, memory-less players) in order to reduce
the search space and come up with a reasonable fix. This work is orthogonal to
ours because we do not try to repair sequential executions of a program.

For concurrent programs, a majority of the work uses dynamic analysis to
repair bugs like atomicity violations [11,12]. For instance, Recon [12] uses test
runs to locate bugs and then uses statistical analysis over these runs to infer a fix.
Being dynamic in nature allows these techniques to scale, but the quality of the
solution is dependant on the coverage of the test runs. On the other hand, our
approach uses static analysis and is capable of providing soundness guarantees
for the fix. Moreover, our notion of a bug is an assertion failure, not notions like
data-races and atomicity violations. Thus, our approach does not get distracted
by benign (and intended) data-races and atomicity violations.

Some techniques require user annotations to infer necessary synchro-
nization. For example, the approach described in [20] infers synchronization once a user
annotates sets of fields, indicating existence of a consistency property within
members of each set.

A quantitative approach to synthesize synchronization has been proposed
in [21]. This work tries to optimize the synthesis of synchronization with respect
to a performance model. Though this work provides correctness as well as perfor-
mance guarantees about the fix, it only works for finite-state programs, making
its use very limited.

We now discuss two pieces of work that are most similar to ours. First is
Wypiwyg (What-You-Prove-Is-What-You-Get) [3], which takes a correct se-
quential library and then synthesizes synchronization (in the form of locks) to
make sure that the library functions correctly even in the presence of a concur-
rent client. Their idea is to take the proof-of-correctness under a sequential client
and then construct synchronization to preserve the same proof even under a con-
current client. This approach contrasts with ours in the following ways: First,
AtomicInf relies on a bug-finding tool, not necessarily ones that can produce
a proof of correctness. Second, AtomicInf guarantees to find the smallest fix
(under atomic sections) irrespective of the underneath verification tool, whereas
the quality of the solution in Wypiwyg depends completely on the quality of
the proof produced—the more modular the proof, the better the synchroniza-
tion inferred. We ran AtomicInf on the benchmarks used by Wypiwyg. Both
approaches inferred the ideal synchronization. However, it is not possible to com-
pare the running times because Wypiwyg used a manually-constructed proof
of correctness for some benchmarks.

The work by Vechev et al. [22] is also very similar to ours. The goal of
their work was to exhibit the power of abstraction-refinement for synthesizing
synchronization using strong atomic blocks. We recast their approach to our
setting in Section 4.1 and then show that our technique (Section 4.2) is more
efficient. Moreover, their work did not address inferring synchronization under weak atomicity.

3 Preliminaries

This section sets up the program syntax used in the rest of the paper and the problem definition. Because we want to control context switching in the program, we assume a co-operative model of concurrency where a program is only allowed to take a context switch at a special **yield** instruction. We write programs using C syntax, extended with the following constructs.

**yield** : The program can context switch only at this statement.

**assume**(*e*) : If the expression *e* evaluates to **false** then the program blocks, otherwise it continues to the next statement.

**axiom**(*e*) : This statement is similar to having **assume**(*e*) at all points in the program. We use **axiom** to insert global invariants into a program.

**satomic**(*stmt*) : This specifies a strong atomic region. *stmt* is executed atomically, i.e., no context switches are allowed while executing *stmt*.

**watomic**(*stmt*) : This specifies weak atomic region. *stmt* is executed in isolation with respect to all other **watomic** blocks. In other words, the execution of *stmt* can not begin if some other thread is executing inside a **watomic** block.

**async** *m()* : This construct spawns a thread which executes method *m()*(). It also returns a handle of the thread created.

**join**(*tid*) : This statement waits for the thread, represented by its handle *tid*, to terminate.

Using the co-operative model of concurrency is not restrictive. Given a multi-threaded program *P*, one can insert **yield** instructions before any instruction that accesses a shared memory location, and also as the first instruction of a thread. The resulting program, under co-operative semantics, is equivalent to *P*. For example, the left side of Fig. 3 shows how the **transfer** method of Fig. 1 is instrumented for co-operative semantics. (For simplicity, we assume that each line of code executes atomically.) Based on this model, we define the notion of a **minimum fix** as follows.

**Definition 1.** A **minimum fix** for a program *P* is one which encloses the least number of **yield** statements under strong or weak atomic blocks.

Sections 4 and 5 address problems of finding a fix under strong or weak atomicity semantics respectively.

Once we have a fix under the co-operative model, we map the fix back to one in the multi-threaded model. Let *Y* be the set of yield instructions that need to be protected by an atomic block, and let *S* be the set of original program locations where these instructions were inserted. Next, we say that two statements *stmt*₁ and *stmt*₂ are **connected** if there is a path from *stmt*₁ to *stmt*₂ or from *stmt*₂ to *stmt*₁ in the control flow graph of the program, such that this path does not
pass through any program point \( p \notin S \). We compute such maximally connected components within the CFG and output it as the atomic blocks. These regions may not be lexically scoped. One way to make them lexically scoped is to consider the set of statements that falls between the dominator and the postdominator of the maximally connected component found earlier. It is a matter of choice whether to output a maximally connected component as a region or augment it to make it lexically scoped.

**Limitations** Although our algorithms guarantee to find the least number of program points to protect in a fix, the process of actually reporting atomic blocks may lose this guarantee; finding lexically-scoped blocks can force us to include other program points in the atomic blocks. However, this is not a major limitation. ATOMICINF also reports the collection of program points and it is usually easy to manually infer the desired fix from this collection of points.

Another limitation is that, in general, the fix inferred by ATOMICINF can only guarantee correctness with respect to safety properties. *It cannot handle liveness properties.* This limitation shows up when the input program itself has some synchronization. Then, imposing the fix inferred by ATOMICINF can lead to deadlocks. For instance, if the program uses flag-based synchronization via a loop: `while(!flag) { }`, (i.e., a thread waits for some other thread to set flag to true), then disabling context switches within the body of this loop can cause a deadlock. We circumvent this problem by never including yield instructions that are meant for synchronization in our fix. This is done partly automatic: yield statements before synchronization operations such as locking routines, and just after an async are excluded from the fix; and partly manual: a user annotates explicit yield points inside shared-memory based synchronization operations, which are also excluded from the fix. We leave a more detailed study for fixing liveness properties as future work.

### 4 Strong Atomicity Inference

Our first step is to gain control over context switching in the program. We do this by introducing a fresh Boolean constant for each yield instruction (except ones excluded because of synchronization—see Section 3), and then guard the yields using this constant as shown in Fig. 3. Let \( CSG \) be the set of Boolean constants introduced this way. Forcing a Boolean constant \( cs_i \in CSG \) to be \( false \) will prevent the context switch from happening at corresponding yield point. For example, in Fig. 3 if we want src->amount to be decremented atomically, we add `axiom(cs_3 == false)` to the program. We also use these Boolean constants to identify the location of a yield instruction.

Given a formula \( \phi \) over \( CSG \), let \( \langle P, \phi \rangle \) be the program \( P \) extended with the statement `axiom(\phi)`. If \( S \subseteq CSG \), then let `disable(S) = \bigwedge_{cs_i \in S} \neg cs_i`. Our goal is to find the smallest set \( S \) such that \( \langle P, disable(S) \rangle \) is a correct program.

For a trace \( t \), let \( CS(t) \subseteq CSG \) be the set of Boolean constants corresponding to the context switches taken in \( t \). Note that \( CS(t) \) cannot be empty when \( t \) is
an error trace of a program without sequential bugs. As previously noted, we assume that the program does not have any sequential bugs. Let $\text{BTraces}(P)$ be the (possibly infinite) set of all error traces of program $P$. Let $\text{CSTraces}(P)$ be $\{\text{CS}(t) \mid t \in \text{BTraces}(P)\}$. Thus, $\text{CSTraces} \subseteq (\mathcal{P}(\text{CSG}) \setminus \{\emptyset\})$, where $\mathcal{P}$ denotes the power set of a given set. Since $\text{CSG}$ is finite, $\text{CSTraces}$ will be finite as well.

For a program $P$, a valid fix is one that rules out all traces in $\text{BTraces}(P)$. To disallow a trace $t$, it is sufficient to disable any one of the context switches taken by $t$. Thus, a fix for $P$ is to disable a set of context switches $S$ such that $S$ is a hitting set of $\text{CSTraces}(P)$. And the smallest fix is a minimum hitting set (MHS) of $\text{CSTraces}(P)$. Note that MHS of any collection of sets need not be unique.

**Definition 2.** Given a set $U$ and a collection of sets $C \subseteq \mathcal{P}(U) \setminus \{\emptyset\}$, a set $H \subseteq U$ is a hitting set of $C$ if $\forall S_i \in C \ S_i \cap H \neq \emptyset$. Furthermore, $H$ is called a minimum hitting set (MHS) if $C$ does not have a smaller hitting set.

Finding an MHS is NP-complete, but for the problem instances that we generate, it is usually quite easy to find an MHS.

### 4.1 A First Approach [22]

Alg. 1 describes an initial approach for finding the smallest set $S$ such that $(P, \text{disable}(S))$ is correct. This approach is inspired from the work of Vechev et al. [22]. Let VERIFTOOL be a verification tool. Given a program, VERIFTOOL($P$) returns $\text{Bug}(t)$ if $P$ has a bug and the error trace is $t$, else it returns $\text{Correct}$.

Alg. 1 iteratively (lines 5-10) finds an error trace $t$ and stores the set of context switches taken by it in $C$. Then $\phi$ is modified to make sure that the
Algorithm 1 Minimum Hitting Set Solution

1: **input:** Concurrent program \( P \) instrumented with Boolean guards for yields.
2: **output:** Set \( S \) of context switches, such that \( \langle P, \text{disable}(S) \rangle \) is correct.
3: \( \phi := \text{true} \)
4: \( C := \emptyset \)
5: **loop**
6: \( \text{res} := \text{VerifTool}(\langle P, \phi \rangle) \)
7: if \( \text{res} == \text{Correct} \) then
8: break
9: end if
10: let \( \text{Bug}(t) = \text{res} \)
11: if \( \text{CS}(t) == \emptyset \) then
12: throw exception("Program has a sequential bug \( t \)"")
13: end if
14: \( \phi := \phi \wedge \left( \bigvee_{c \in \text{CS}(t)} \neg c \right) \)
15: \( C := C \cup \{ \text{CS}(t) \} \)
16: **end loop**
17: return \( \text{MHS}(C) \)

Algorithm 2 Optimized Minimum Hitting Set Solution

1: **input:** Concurrent program \( P \) instrumented with Boolean guards for yields.
2: **output:** Set \( S \) of context switches, such that \( \langle P, \text{disable}(S) \rangle \) is correct.
3: \( \phi := \text{true} \)
4: \( C := \emptyset \)
5: **loop**
6: \( \text{res} := \text{VerifTool}(\langle P, \phi \rangle) \)
7: if \( \text{res} == \text{Correct} \) then
8: break
9: end if
10: let \( \text{Bug}(t) = \text{res} \)
11: if \( \text{CS}(t) == \emptyset \) then
12: throw exception("Program has a sequential bug \( t \)"")
13: end if
14: \( \phi := \text{disable}(\text{MHS}(C)) \)
15: \( C := C \cup \{ \text{CS}(t) \} \)
16: **end loop**
17: return \( \text{MHS}(C) \)

As mentioned in Section 2, this algorithm is not very efficient. Consider the code snippet shown in Fig. 4. Suppose there are two threads executing this code. The first thread executes the code on the left and the second thread executes the code on the right. The program fails whenever the statement \( x = 5 \) gets interleaved between statements \( x = 10 \) and the assertion. Assignments to \( \text{tmp} \) are redundant but they introduce extra yield points.

When we run Alg. 1 on this code, we can get error traces that first execute \( x = 10 \), then context switch at \( \text{cs1} \), then execute \( x = 5 \) and some part of the second thread, context switch at \( \text{cs}_i \) (for \( 3 \leq i \leq N \)), and then fail the assertion. There can be \( N - 1 \) such traces. Thus, Alg. 1 will potentially make \( N - 1 \) calls to VerifTool. While the fix is to disable just \( \text{cs1} \), the number of verification calls made by this approach is proportional to the size of the program.

Fig. 4: A code snippet.

```bash
if(cs1) yield();
x = 10;
if(cs2) yield();
assert(x == 10);
if(cs1) yield();
tmp = 1;
tmp = 1;
... if(csN) yield();
tmp = 1;
```
4.2 Our Approach

Alg. 2 improves the previous algorithm by being more efficient when the size of the solution is small. The main difference is on line 15. It computes a proposed solution by looking at all previous traces. To see why this is an improvement, let us again consider the program in Fig. 4. Suppose the first trace takes context switches $S_1 = \{cs_1, cs_5\}$. Then $C = \{S_1\}$ and it has two possible choices of MHSs. Suppose (unluckily) we pick $MHS(C)$ as $\{cs_5\}$. Then $\phi$ disables $cs_5$. The VerifTool call will return another error trace passing through, say, $S_2 = \{cs_1, cs_8\}$ (note that all error traces have to take $cs_1$). Then $C = \{S_1, S_2\}$ and has exactly one MHS, which is $\{cs_1\}$. Thus, we converge to the desired solution in just two queries, independent of $N$. Furthermore, the constraints $\phi$ added to the program $P$ are much simpler than the ones added by Alg. 1, making the job of the verifier easier.

**Theorem 1.** Given a program $P$ with no sequential bugs, Alg. 1 and Alg. 2 compute a minimum hitting set of $\text{CSTraces}(P)$.

*Proof.* Let $m$ be the MHS of $\text{CSTraces}$. Each of the algorithms returns an MHS over some subset of $\text{CSTraces}$. Let $C_i$ be the subset used by Alg. $i$ and let $m_i$ be its MHS. Both $m_1$ and $m_2$ are valid fixes because VerifTool eventually returns CORRECT. Thus, both are hitting sets of $\text{CSTraces}$. Because $m$ is an MHS of $\text{CSTraces}$, it must be a hitting set of $C_1$. This implies $|m_1| \leq |m|$. Thus, $m_1$ is an MHS of $\text{CSTraces}$. Same argument applies for $m_2$.

**Performance comparison between Alg. 1 and Alg. 2:** A direct theoretical comparison between the running times of Alg. 1 and Alg. 2 is difficult because of inherent non-determinism in these algorithms. In particular, the verification tool may return any arbitrary buggy trace in the program fed to it, making it possible for any of Alg. 1 and Alg. 2 to get "lucky" and converge to a fix faster. However, we can show that if both algorithms witness the same set of traces, then Alg. 2 is never worse than Alg. 1.

Let $\phi_{alg1}$ and $\phi_{alg2}$ denote the constraints generated by Alg. 1 and Alg. 2 respectively, on lines 14 and 15. Further, suppose that the first $n$ iterations of the algorithms witness the same traces $t_1, \ldots, t_n$. Then it must be that $\phi_{alg2}$ is stronger than $\phi_{alg1}$. For every trace $t_i$, $\phi_{alg1}$ has a clause $\bigwedge_{cs \in \text{CS}(t_i)} \neg cs$. On the other hand, $\phi_{alg2}$ has a clause with a single literal $\neg cs'$, where, $cs'$ is the context switch taken by $t_i$ and is a part of an MHS computed by it. Then $\phi_{alg2} \to \phi_{alg1}$ follows from $a \to a \lor b$ (for each clause corresponding to a trace) as well as $a \to b \land c \to d \Rightarrow a \land b \to c \land d$ (conjunction of clauses from all the traces). Consequently, if Alg. 1 terminates in the $n + 1$st iteration, then so will Alg. 2. Our experiments(Section 3) show the superiority of Alg. 2 in practice.

5 Weak atomicity inference

Computing a fix using weak atomicity is harder because it doesn’t directly allow us to disable context switches. We set up some terminology first.
Definition 3. Given a trace $t$ and a context switch $cs$ taken by $t$, let $T$ be the thread that was executing when $cs$ was taken. Then the lifespan of $cs$ in $t$ is defined as the set of all instructions (or program points) on $t$ after $cs$ but before $T$ got control back. In other words, the lifespan of a context switch is the contiguous sub-trace between the two instructions of the same thread that surround the context switch.

For example, suppose $t = [a_1; a_2; a_3; b_1; b_2; c_1; c_2; b_3; a_4]$, where $a_i$ denote instructions of thread 1, $b_i$ denote instructions of thread 2, and $c_i$ denote instructions of thread 3. Then the lifespan of the context switch at $a_3$ is $\{b_1, b_2, c_1, c_2, b_3\}$.

The way to rule out an error trace $t$ using weak atomicity is to pick two yield instruction $y_1$ and $y_2$ on the trace such that: (1) $y_1$ appears before $y_2$; (2) $t$ context switches at $y_1$ and (3) the lifespan of the context switch at $y_1$ includes $y_2$. In this case, we say that $y_2$ conflicts with $y_1$. Then including both $y_1$ and $y_2$ in a weak atomic block will render $t$ infeasible. Moreover, this is the only way to disable a trace using weak atomic blocks (without inserting or deleting extra code). In contrast, for strong atomicity, we only had to look at $y_1$. Thus, weak atomicity forces us to identify the conflict between threads.

As before, we introduce a Boolean constant for every yield instruction in the program. Furthermore, we introduce a global Boolean variable $lock$ that is initialized to $false$. If $cs$ is the Boolean constant associated with a yield, then we transform it as follows:

$$\text{yield}; \Rightarrow \begin{cases} \text{if}(\neg cs) \{ \text{assume lock} = false; \text{lock} = true; \} \\ \text{yield}; \\ \text{if}(\neg cs) \{ \text{lock} = false; \} \end{cases}$$

This way, setting a Boolean constant to $false$ is as if the corresponding $\text{yield}$ is included in a weak atomic block.

For a trace $t$, let $\text{WCS}(t)$ be the set of $(cs_1, cs_2)$ pairs such that $cs_1$ corresponds to a yield instruction $y_1$, and $y_2$ conflicts with $y_1$. Let $\text{WCSTRACES}(P)$ be $\{\text{WCS}(t) \mid t \in \text{BTRACES}(P)\}$. The smallest solution is given by the MHS of $\text{WCSTRACES}(P)$. If this set is $W$, then the following set of yields need to be protected by a weak atomic block: $\{y \mid (y_1, y) \in W \text{ or } (y, y_2) \in W\}$.

We can now set up our algorithm in a similar fashion to Alg. 2. However, we now have to gather pairs of instructions, which can lead to a large number
of iterations. So we make use of a crucial optimization: First, we compute the strong atomicity solution $S \subseteq CSG$ for the program. Next, we only attempt to find the smallest extension of this solution that will fix the program using weak atomicity blocks. This is done as follows: For a trace $t$, instead of using $WCS(t)$, we use $WCS(t, S) \equiv \{ cs_2 \mid \exists cs_1 \in S : (cs_1, cs_2) \in WCS(t)\}$. Note that for an error trace $t$, if $WCS(t)$ is not empty then neither is $WCS(t, S)$ because we know that some context switch taken by $t$ belongs in $S$. Thus, we only look for conflicts with context switches in the strong atomicity fix. Because $WCS(t, S)$ is a subset of $CSG$, we are again back to iterating over $CSG$ rather than $CSG \times CSG$. Alg. 3 formalizes this description.

The penalty of using this optimization is that we do not guarantee the smallest fix, however, we do guarantee the smallest extension to the strong atomicity fix, and in our experiments we always obtained the smallest fix possible.

6 Implementation and Experiments

We have implemented Algs. 2 and 5 in a tool called AtomicInf. We use Poirot \cite{10,16} as the underlying verification tool. Poirot is really a bug-finding tool; it searches over all behaviors up to a bounded number of context switches, thus, it cannot prove the absence of bugs. In this case, the fix returned by AtomicInf is correct only up to the capabilities of Poirot. In our experiments, we manually verified that the the computed fixes were sound. In principle, we could have used a true verification tool like Threader \cite{7} inside AtomicInf to obtain sound fixes.

Results We evaluate the effect of changing various parameters on the performance of Algs. 1, 2, and 5. Consider the parameterized program shown in Fig. 5. It has two threads: the first executes the code on the left and the second thread executes the code on the right. The program has three parameters $p_1, p_2,$ and $p_3$ that control the program size. Note that the strong atomicity fix is to enclose the entire body of the first thread in an atomic block. Thus, the size of the strong atomicity fix is $p_1 + 1$ (the number of yields inside this block of code). The size of weak atomicity fix is $p_1 + p_2$ because all of the assignments to $x$ in the second thread must be put inside a weak atomic block as well. The parameter $p_3$ controls the number of irrelevant assignments to shared variables. Results are shown in Fig. 6. Here, $\#CS$ is the number of yield instructions inserted in the program, $\#Q$ indicates the number of queries made to Poirot and the last column indicates running time in seconds. Compare Fig. 6(a) with Fig. 6(c). As expected, Alg. 1 requires more calls to Poirot as the size of the program increases. However, the number of calls made by Alg. 2 remains constant irrespective of the program size. Figs. 6(b) and 6(d) show that the number of queries required by Algs. 2 and 5 increases almost linearly as the size of the solution increases. Here, (W) in columns indicates the numbers for weak atomicity fix.

Next, we ran AtomicInf on various benchmarks gathered from previous work. The results are shown in Tab. 1. In the table, LOC is lines of code, $\#CS$
Fig. 5: A parameterized program with two shared variables: x and y. Here, \([st]n\) denotes the statement \(st\) repeated \(n\) times.

\[
\begin{align*}
&x = 10; \\
&[y = 1;]p_1 \\
&assert(x == 10); \\
&[x = 1;]p_2 \\
&[y = 1;]p_3
\end{align*}
\]

Fig. 6: Effects of changing various parameters of the program in Fig. 5.

is the number yield instructions inserted in the program, Sol Size is the number of program points as part of the computed fix, \#Queries is the number of times POIROT was called and the last column is the running time in seconds. The sub-columns S1 and S2 indicates results for strong atomicity by Alg. 1 and Alg. 2 respectively. W indicates weak atomicity results obtained by running Alg. 2 followed by Alg. 3. Numbers in bold indicates the better results amongst Alg. 1 and Alg. 2. Against each benchmark, we refer to the paper from which it was obtained. Here, banking in paper is the running example used in Fig. 1. Both the algorithms converged to the same solution for strong atomicity for all the examples. On the average Alg. 1 takes 20% more queries and 74% more time as compared to Alg. 2. If we discount for the outlier benchmark "Bank Account", Alg. 1 requires twice the number of queries on the average. As mentioned in Section 4.2 non-determinism plays a role as the two algorithms witness different set of traces and takes different amount of time. It is important to note that the most expensive operation in terms of time is a call to VERIFTOOL. We have observed that most of the time is spent inside the subroutine VERIFTOOL. Compared to this, the time consumed in computing MHS is negligible. For all of the examples, we manually inspected as well as cross verified with the papers from which the benchmarks were taken. We found the quality of the solution proposed by ATOMICINF to be the smallest and precise. On the other hand, for programs
Table 1: Results of running ATOMICINF on a number of program snippets with published concurrency bugs.

| Example                 | LOC | #CS | Sol Size | # Queries | Time(sec) |
|-------------------------|-----|-----|----------|-----------|-----------|
| banking_inpaper(fig.1)  | 62  | 22  | 2        | 22        | 9         | 35.9      |
| banking_inpaper_corpus  | 58  | 23  | 3        | 12        | 7         | 16.6      |
| apachex [19]            | 64  | 10  | 2        | 2         | 4         | 4.6       |
| mozilla [11]            | 68  | 12  | 3        | 4         | 4         | 3.2       |
| apache1 [19]            | 64  | 11  | 2        | 2         | 2         | 2.9       |
| mozilla1 [11]           | 64  | 7   | 2        | 2         | 4         | 2.9       |
| mysql [15]              | 70  | 13  | 4        | 6         | 6         | 9.4       |
| banking [23]            | 231 | 52  | 4        | 13        | 5         | 298.5     |
| defrag [22]             | 142 | 37  | 2        | 2         | 6         | 475.5     |
| doubleLockQueue [14]    | 144 | 29  | 4        | 8         | 5         | 321.4     |
| interceptMessagePane [12]| 211| 68  | 2        | 3         | 12        | 85.3      |
| parkingBufferBoot [12]  | 219 | 46  | 1        | 2         | 3         | 89.6      |
| BankAccount [12]        | 149 | 32  | 11       | 4         | 13        | 16.9      |
| CircularList [12]       | 149 | 32  | 11       | 4         | 13        | 16.9      |
| StringBuffer [12]       | 139 | 29  | 8        | 9         | 9         | 232.8     |
| logProcessNSweep [12]   | 149 | 32  | 11       | 4         | 13        | 16.9      |
| compute [3]             | 61  | 7   | 2        | 2         | 6         | 9.7       |
| average [3]             | 69  | 14  | 3        | 24        | 24        | 16.3      |
| increment [3]           | 24  | 9   | 1        | 2         | 3         | 14.7      |
| nonDetRet [3]           | 49  | 26  | 3        | 5         | 8         | 43.1      |

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