Multi-Variant Execution of Parallel Programs

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Abstract—Multi-Variant Execution Environments (MVEEs) are a promising technique to protect software against memory corruption attacks. They transparently execute multiple, diversified variants (often referred to as replicae) of the software receiving the same inputs. By enforcing and monitoring the lock-step execution of the replicae’s system calls, and by deploying diversity techniques that prevent an attacker from simultaneously compromising multiple replicae, MVEEs can block attacks before they succeed. Existing MVEEs cannot handle non-trivial multi-threaded programs because their undeterministic behavior introduces benign system call inconsistencies in the replicae, which trigger false positive detections and deadlocks in the MVEEs. This paper for the first time extends the generality of MVEEs to protect multi-threaded software by means of secure and efficient synchronization replication agents. On the PARSEC 2.1 parallel benchmarks running with four worker threads, our prototype MVEE incurs a run-time overhead of only 1.32x.

Index Terms—replication, multi-threading, monitoring, determinism, developer effort, performance overhead

1 MOTIVATION

Memory corruption vulnerabilities have been used to undermine the security of computer systems for at least four decades. To this day, new attacks appear shortly after new mitigations are introduced and long before they can be put into practice. A clear trend is that transparent mitigations with low overheads are able to enter practice. Address Space Layout Randomization [1], Stack Cookies [2] and Data Execution Prevention [3] are technologies that meet the high bar to adoption. However, both the broadly deployed defenses and many emerging ones have already been bypassed [4], [5], [6], [7], [8], [9], [10], [11]. Solutions that provide much stronger protection are less transparent and involve more overhead. They include Software-based Fault Isolation [12], Control-Flow Integrity [13], Code-Pointer Integrity [14], and Code Diversification [15].

The widespread availability of multi-core processors has made Multi-Variant Execution Environments (MVEEs) an increasingly attractive solution. MVEEs monitor multiple instances, known as replicae, of the same program running in parallel. When the MVEE detects inconsistencies in the replicae’s I/O behavior, it considers those as symptoms of an ongoing attack and it terminates the replicae’s execution to prevent harm.

The security guarantees provided by an MVEE build on three properties: (i) isolation of the monitor from the replicae; (ii) monitored lock-step system call execution; and (iii) diversification of the replicae. The isolation of the monitor by means of hardware-enforced boundaries is achieved by implementing it in kernel-space or in a separate user-space process. All system calls invoked in the replicae are monitored, executed in lock-step, and only allowed to execute when all replicae invoke the same system calls with consistent inputs. This allows the monitor to detect when a single replica is compromised and to halt its execution. However, this also implies that all replicae only progress at the speed of the slowest one. Furthermore, it implies that the monitor cannot look ahead at future system calls when deciding if a specific system call invocation should be allowed. This places strict constraints on the system calls in the replicae. Foremost, they need to occur in the same order in all replicae.

Those replicae are diversified instances of the same program. Several techniques are suitable to generate the diversified replicae, including Address Space Partitioning [16], Reverse Stack Growth [22], System Call Number Randomization [23] and Disjoint Code Layout [24]. All of these techniques aim to introduce enough diversity to cause detectable divergences in the system call behavior of replicae under attack, while maintaining consistent system call behavior if the inputs are benign. This task is complicated by the fact that the system call behavior does not solely depend on explicit, user-provided program inputs, but might also depend on implicit inputs.

We prevent benign system call consistency violations that result from implicit inputs, such as randomized virtual addresses that affect the system call behavior through address-dependent computations, by using implicit-input replication agents [21]. These agents log the implicit inputs in one replica, the master, and replicate those inputs in the other replicae, the slaves, thus restoring system call consistency. Together with the system calls, the points at which these agents intervene form the so-called rendez-vous points (RVPs) of the replicae.

In this work, we focus on a big weakness of existing secure MVEEs with respect to system call consistency: secure replication of non-deterministic multi-threaded applications. In real-life multi-threaded programs, even the security-sensitive system calls that should be monitored most strictly often differ between replicae as a result of their non-deterministic scheduling. Because of the lock-step system call execution and monitoring requirement, security-oriented MVEEs cannot tolerate those divergences.

Two broad classes of techniques could potentially alleviate this problem. First, a Deterministic Multi-Threading (DMT) system can be embedded in the program to enforce a fixed thread schedule in all replicae [25], [26], [27], [28], [29].
In the context of an MVEE, however, all existing DMT systems fall short. To establish a deterministic schedule, they all rely, in one way or another, on a token. Some systems only allow threads to pass this token when they invoke a synchronization operation. This approach is incompatible with threads that deliberately wait in an infinite loop for an event such as the delivery of a signal to trigger because such threads may never invoke a synchronization operation. Other systems allow threads to pass their token when they have executed a certain number of instructions. Such systems cannot tolerate variations in the program execution and are therefore incompatible with most code diversification techniques as well as the implicit-input replication agents.

Alternatively, we can accept non-determinism and require only that all replica execute in the same non-deterministic order. Online Record/Replay (R+R) systems can provide this guarantee by logging the execution in one replica and replaying it in the other replica. R+R systems are less sensitive to variations in the program execution, which we typically see with diversified replica. But in order to use them in a security-oriented MVEE, they need to be adapted to become address-agnostic and to support programs that use ad hoc (i.e. non-standardized) synchronization primitives or lock-free algorithms. Furthermore, for embedding an R+R system in a security-oriented MVEE, we need to ensure that any new functionality introduced in the replica must be neutral with respect to the RVPs, and we need to secure the communication channel that is used to convey the information about the recorded execution from the master replica to the slave replica.

Our paper makes four contributions. First, we present four R+R-based synchronization replication agents that record synchronization operations in a master replica and replay them in the slaves, thus ensuring system call consistency. The agents are address-agnostic and system-call-neutral, and hence compatible with existing secure MVEEs and implicit-input replication agents. The most efficient agent communicates over a channel that is secured against malicious communication by attackers.

Second, we present a practical strategy to extend our R+R-based systems to support ad hoc and lock-free synchronization, which we typically see in many low-level libraries.

Third, we report how we integrated our replication agents into GNU’s glibc and how we applied the aforementioned strategy to four commonly used system libraries: GNU’s libthread, libstdc++ and libgomp. This integration enables support for data race free C and C++ programs that use pthread and/or OpenMP programming models.

Finally, we extensively evaluate the run-time performance of our replication agents, the implementation effort that went into their integration into the aforementioned libraries, and the security of the proposed features.

2 Replication of Multi-Threaded Programs

The techniques presented in this paper build on GHUMVEE, an existing security-oriented MVEE. To monitor the replica, GHUMVEE uses the ptrace and process_vm_* Linux APIs. As the use of these APIs involves context switching, they introduce significant latencies in the interaction between the monitor and the replica. This makes them unacceptable for replicating synchronization events, which occur frequently in many programs and which are often handled entirely in user space in the original programs to optimize performance. For example, we observed gcalctool, a simple calculator from the GNOME desktop environment performing 1.8M futex operations during its 400 ms initialization, almost all of which were uncontended and hence handled in user space. Interposing all those operations with system calls and ptrace made the initialization time grow to over 370 seconds, a slowdown with a factor 925!

To avoid such an unacceptable overhead, our alternative solution consists of a synchronization replication agent that replicates all synchronization events entirely in user space.

2.1 Synchronization Replication Agent

We enforce an equivalent execution in all replica by injecting a synchronization replication agent into their address space, as shown in Figure 1. At run time this agent forces the master replica to capture the order of all inter-thread synchronization operations, hereafter referred to as sync ops. The agent logs the captured order in a circular, shared buffer that is visible to all replica. This buffer is mapped with read/write permission in the master replica and with read-only permission, and at different addresses, in the slave replica. In the slave replica, the agent uses the captured order to enforce an equivalent replay of sync ops.

To capture the sync op execution order, we wrap them in a small critical section at the source code level. Within the critical section we first log information about the sync op in the first available slot of the buffer and then perform the original op. Depending on the agent we use, the information about the sync op consists of the thread ID, the memory word that was affected by the op, and the op’s type.

The replication agent must be available to the entire program, including any loaded shared libraries. So we chose to implement the agent in glibc, at the lowest possible level in the user-mode portion of the software stack where it is exposed to the program itself and to all shared libraries.

The same agent is used in the master and slave replica. Identical instances of glibc are loaded into the master and slave replicae when they are launched, though they might be loaded different addresses in each replica. When an instance is later invoked at run time, it needs to know whether it is invoked in a master or slave replicae, to either capture or replay the sync op order. We therefore dynamically initialize the agent in each replica. Soon after a replica is launched by the MVEE monitor, its agent invokes a system call that is intercepted by the monitor. Through this system call, the
more, because a secure MVEE needs to enforce lock-step execution on the replica, we need to replicate information actively and with minimal delay to avoid that slave replica delay the master replica too much. We therefore implement replication agents that make the recorded information visible immediately, rather than broadcasting it periodically.

These two design decisions have far-reaching consequences. First, the RVP-neutrality constraint prevents us from using dynamic memory allocators, because those use sync ops to coordinate multi-threaded access to the memory, and introduce system call RVPs to allocate additional memory pages and to change protection flags.

Second, since we want any information to be visible immediately in the slave replicae, the agent cannot perform any post-processing on the recorded information. This prevents us from compressing the recorded information to reduce our agents’ memory bandwidth requirements.

Third, our agent must support diversified replicae. It can therefore not assume that the master and slave replicae are fully identical. For example, the same mutex might be found at different addresses in the different replicae. Consequently, the recording side of the replication agent must record its information in a manner that is address-agnostic.

### 2.2 Replication Strategies

To replay the sync ops in the slave replicae, several approaches are available that trade CPU cycles off against memory pressure. We have implemented three replication strategies that meet the aforementioned constraints.

#### 2.2.1 Total-order replication agent

Our total-order (TO) replication agent replays all sync ops in the exact same order in which they happened in the master replica. Figure 2(a) shows two threads that execute under GHUMVEE’s control. In the master replica, thread M1 first enters and leaves a critical section protected by lock A at times $t_0$ and $t_1$ resp. At those times, the wrappers of the corresponding sync ops log the activities of thread M1 in the replication buffer. Next, thread M2 in the master replica enters and leaves a critical section protected by lock B at times $t_2$ and $t_3$ resp. These events are also logged in order in the buffer. Right after $t_3$, the buffer holds the contents indicated in the figure. Time stamps to the left and right of the buffer mark the time the buffer elements are produced and consumed resp. The arrows on the left and right denote the position of the producer and the consumer pointers resp. right after $t_3$.

In the slave replica thread S2, corresponding to M2 in the master replica, reaches the critical section protected by lock B first, at time $t_4$. At that time, the first element in the buffer indicates that synchronization events in the master replica occurred in thread M1 first, so thread S2 is stalled in the wrapper of the sync op in enter_sec. Only after the first two elements in the buffer are consumed in thread S1 at times $t_5$ and $t_6$, can thread S2 continue executing. Thus, even though the two critical sections protected by locks A and B are unrelated, thread S2 is forced to stall until thread S1 has replayed the operations performed by thread M1.

This agent is trivial to implement, but not very efficient: The lack of lookahead by consumers introduces unnecessary stalls as indicated by the red bar in Figure 2(a).
2.2.2 Partial-order replication agent

Our partial-order (PO) replication agent is more efficient. It only enforces a total order on dependent sync ops. This agent may replay independent sync ops in any order, as long as it preserves sequential consistency within the thread. The PO agent is more complex and introduces more memory pressure because the agents in the slave threads have to scan a window in the buffer to look ahead. However, it typically introduces much less stalling and generally outperforms the TO agent. In Figure 2(b), we see the exact same order of events as in Figure 2(a) until \( t_4 \). This time, however, thread \( S_2 \) may enter the critical section without delay at \( t_3 \) because the enter_seq operation does not depend on either of the operations that preceded it in the recorded total order.

Conceptually, there are significant similarities between this agent and online R+R techniques such as LSA \( [36] \) and offline R+R techniques such as RecPlay \( [39] \). However, our agent captures events at a finer granularity of sync ops instead of pthread-based synchronization operations. Furthermore, our agent is fully RVP-neutral. These are relatively minor differences, however, as techniques like LSA can be adapted to capture at a lower granularity and to use only statically allocated memory. A more fundamental difference is that our agent also supports diversified replica. With queue projection, LSA discards the per-thread order of synchronization operations and only maintains the per-variable order. With diversified replica, the same logical variable might be stored at different addresses in different replica. Our agent therefore relies on the per-thread order to determine which logical variable is affected by each synchronization operation.

Although the PO agent eliminates unnecessary stalling, it still suffers from poor scalability. The master replica must safely coordinate access to the circular buffer by determining the next free position in which it can log an operation. If many threads simultaneously log synchronization events, this inevitably leads to read-write sharing on the variable that stores the next free position. A similar problem exists on the slave replica’s side because they must keep track of which data has been consumed. With multiple slave replica, this also leads to high sharing and, consequently, high cache pressure and cache coherency traffic.

2.2.3 Wall-of-clocks replication agent

The above observation led us to the design a third agent. This wall-of-clocks (WoC) agent assigns each distinct memory location that is ever involved in a sync op to a logical clock. These clocks capture “happens-before” relationships between related sync ops \( [40] \). Similar to, e.g., plausible clocks, but without using clock vectors, our clocks only capture the necessary relationships \( [41] \).

In Figure 2(c), lock A stored at address \( &A \) is assigned to clock \( cA \). Lock B is similarly assigned to clock \( cB \).

On the master side, the agent logs the identifier of the logical clock associated with each sync op, as well as that clock’s time. After logging each sync op, the agent increments the logical clock time of the associated clock.

In this agent, the logging is no longer done in a single circular buffer. Instead there is one circular buffer per master thread, such that each buffer has only one producer. In Figure 2(c), master thread M1 only communicates with slave thread \( S_1 \) through buffer 1, whereas thread \( M_2 \) only communicates with thread \( S_2 \) through buffer 2. This design avoids the contention for access to the shared buffers.

Neither the master nor the slave replica need to propagate their current buffer positions to other threads. Furthermore, the master’s logical clocks do not need to be visible to the slaves. The information contained within the circular buffers suffices for the slave replica to replay the same clock increments on their own local copies of each clock.

In Figure 3(c), thread \( M_1 \) first enters a critical section protected by lock A at time \( t_0 \). The agent observes that the current time on logical clock \( cA \) is 0. It records the clock and its time in buffer 1 and increments the clock’s time to 1. At time \( t_1 \), the agent logs the exit from the critical section in buffer 1. This time around, the logical clock time is 1.

A similar situation then unfolds in thread \( M_2 \) at time \( t_2 \). This time though, the critical section is protected by lock B, of which the associated memory location is assigned to clock \( cB \), whose initial time also is 0. This information is logged in circular buffer 2, along with information regarding the exit of the critical section in thread \( M_2 \) at time \( t_3 \). At that point, clock \( cB \) is incremented to 2.

In thread \( M_1 \), a third critical section is entered at time \( t_4 \), which is again protected by lock B. This event involving logical clock \( cB \) is logged in buffer 1 with clock time 2.

On the slave replica’s side, the threads are scheduled differently in our example. There, thread \( S_2 \) reaches a sync op first, at time \( t_5 \). The agent observes in buffer 2 that it must wait until clock \( cB \) reaches time 0. Since this is the initial time on the slave’s copy of that clock, the operation can be executed right away and thread \( S_2 \) will increment the time on its copy of \( cB \) to 1. If we suppose that thread \( S_2 \) is then pre-empted and thread \( S_1 \) gets scheduled, \( S_1 \) will enter and leave the critical section protected by lock A at times \( t_6 \) and \( t_7 \), consuming the first two entries in buffer 1, thereby incrementing the slave copy of clock \( cA \) to 2.

The third operation in thread \( S_1 \) at time \( t_8 \) is the most interesting. In the first replication buffer, the slave agent observes that the sync op to enter a critical section has to wait until its associated logical clock \( cB \) has reached time 2. However, in the slave, that clock’s time was last incremented at time \( t_5 \), i.e., to the value of 1. Thread \( S_1 \) must therefore wait until some other slave thread has incremented the time on \( cB \). This will happen at time \( t_9 \) in thread \( S_2 \). Shortly thereafter, the agent code executing in thread \( S_1 \) will observe that \( cB \) has reached the necessary value, and at time \( t_{10} \) \( S_1 \) will enter its second critical section.

With this WoC, the replication agent only inserts accesses to shared data, and hence coherence traffic, for two reasons. First, it introduces accesses to replication buffers shared between corresponding threads in the master and slave replica. This is a fundamentally unavoidable form of overhead required to replicate the synchronization behavior from the master to the slave replica.

Secondly, the agent inserts accesses to shared clocks whenever multiple threads in the original program were already contending for locks at shared memory locations. While these extra shared accesses in the replication agents still introduce some overhead, we do expect the overhead to scale with the pre-existing resource contention in the orig-
inal application. In other words, if the original application uses contended global locks that decrease the available parallelism, the replication agent will hurt it further. However, if the original application involves a lot of synchronization, but that synchronization is performed using uncontended local locks, the WoC replication agent will not introduce contended traffic within the master or slave replica.

As we will see in Section 3, the WoC agent consistently outperforms the other agents on almost every benchmark. Most importantly, as is the case with plausible clocks in general, the replication will always be correct.

One important remark remains to be made, however. While the WoC agent is certainly the more elegant and more efficient of the three proposed designs, it is not fully optimal. Due to the RVP-neutrality constraint, we cannot dynamically assign each memory location to its own private clock. Instead, we have to pre-allocate a fixed number of clocks statically and we have to assign lock memory locations to one of those clocks based on a hash of their memory address. Because we want to use a cheap hash function, hash collisions are quite likely. Any such collusion results in an m-to-1 mapping between multiple clocks and each clock. In other words, the WoC agent is bound to assign some non-conflicting memory locations to the same logical clock. When this happens, this introduces unnecessary serialization and hence potentially also unnecessary stalls in the slave replica.

Our WoC agent is similar to Respec [37], although it does not share any part of its implementation. It differs from other clock-based techniques, however, in that it does not use thread clocks. Instead, our agent relies solely on the logical clock it assigns to each memory location. In the ideal case, our agent therefore only needs to read and update the value of one clock to replay a synchronization operation. Techniques that rely on Lamport clocks (e.g. ROLT [42]) by contrast need to read and update the values of two clocks: the local thread clock and the synchronization variable’s associated clock. Techniques that rely on vector clocks (e.g. RecPlay [39]) need to read the value of at least n + 1 clocks (with n the number of threads in the program): the local thread clock, the synchronization variable’s clock, and the thread clocks of all other threads. The reason why our agent does not need local thread clocks is that it records into a per-thread buffer, rather than a globally shared buffer. Therefore, a thread clock would never have to be synchronized with other thread clocks, which eliminates the need for such a clock altogether. Furthermore, the fact that our agent assigns each memory location to a statically allocated clock implies that the agent can be applied transparently and that it respects RVP-neutrality.

### 2.3 Secured wall-of-clocks agent

The agents implementing the three replication strategies as discussed in the last sections are not very secure: They forward information through a circular buffer that is shared among all replicas. This buffer easy to locate since all three of these agents store a pointer to it in a thread-local variable. Despite of the code reuse countermeasures we have in place [24], attackers could exploit the fact that an easily locatable communication channel between the replicas exists to set up an attack that can compromise multiple replica.

As the WoC agent outperforms the other two agents on average, we build on that agent to present an alternative, secured design that relies on the hidden buffer array (HBA) shown in Figure 3. This page-sized array stores pointers to hidden buffers. Upon startup, a replica can request that this HBA be allocated by GHUMVEE and subsequently map it into its own address space using the System V IPC API [43]. GHUMVEE intercepts and manipulates this mapping call such that the pointer to the HBA is not returned to the program. At the same time though, GHUMVEE overrides the base address of the replica’s gs segment so that it points to the HBA.

The reason to override this address is that the x86_64 architecture supports addressing of 48-bit (or bigger) pointers and has therefore disabled most of the original x86 segmentation functionality. The gs and fs segment registers may still be used as additional base registers, however, and by consequence all gs or fs-relative memory accesses are still valid. It is extremely uncommon to still find such accesses in x86_64 software, however. Furthermore, x86 processors do not allow user-space instructions to read the segment registers. The gs and fs segments can therefore be used to store pointers that are hidden from the user-space software.

At a fixed offset within the HBA we store a pointer to the agent’s circular buffer. The end result is that the replica must read the pointer to the circular buffer indirectly, through a gs-relative memory access. In assembler, we manually crafted a version of our WoC agent that accesses this pointer in such a way. By storing the pointer to the buffer and any pointers derived from it in a fixed caller-saved general-purpose register, we guarantee that the pointer never leaks to memory and that no function outside the replication agent can observe the pointer value. We further guarantee that (i) the pointer to the buffer is never moved to a different register, (ii) the register is never pushed onto the stack, (iii) the register is cleared before the function returns and (iv) the replication agent does not call any functions while the pointer value is visible.

Since neither gcc, nor LLVM offers syntactic sugar to allow for such properties, we have chosen to implement both of the replication agent’s functions that access the shared buffer in assembly code. The current implementation, which we evaluate in Section 5 totals approximately 150 LoC.

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**Fig. 3: Hidden buffer array access to the replication buffer.**
2.4 Embedding the replication agent

The key challenge in embedding the replication agent into a program is to identify all sync ops. Because we want to wrap these sync ops in the source code itself, we also identify the sync ops at the source code level.

Existing R+R systems, as well as DMT systems that impose weak determinism, only order invocations of pthread-based synchronization functions. That is insufficient for a secure MVEE because glibc and several other low-level libraries implement their own sync ops. A failure to order the sync ops in glibc tends not to affect the user-observable I/O determinism of the program, but it does impact the general system call behavior and hence violates the system call consistency needed in a secure MVEE.

An alternative strategy would be to order all sync ops by wrapping all loads and stores in the program. This would yield system call consistency even in the presence of data races. Ordering individual loads and stores leads to prohibitively high overhead however, as was demonstrated in the context of strong determinism [33]. Moreover, given the range of diversity we need to support in the replication agent that enforces an equivalent order of sync ops in all replicae, we use a partially manual process. First, we disassemble the binary code, and then translate those to source line numbers by means of debug information.

The strategy we propose is most similar to weak determinism systems, but we capture sync ops at a lower level as shown in Figure 4. In existing strong determinism systems, all 11 memory operations need to be ordered. (Lines 3, 9, and 10 each involve two memory operations.) With existing systems of weak determinism, only the two (standard synchronization) operations on the mutex on lines 2 and 5 are ordered. In our solution, we wrap both the standard operations on the mutexes as well as the three ad hoc synchronization events on lines 7, 11 and 13. The latter one translates into a LOCK SUB instruction on the x86 architecture, and atomically sets the zero flag (ZF).

To identify the source lines to be wrapped, we first identify the sync ops in the binary code, and then translate those to source line numbers by means of debug information.

The relevant sync ops in the binary code come in three categories. First, any instructions that the programmer explicitly marks as atomic are sync ops. On the x86 architecture, this includes all instructions with an explicit LOCK prefix, as well as XCHG instructions with an implicit LOCK prefix [44]. LOCK CMPXCHG is an example. Second, any store operation (e.g., for a C assignment to a dereferenced pointer or volatile variable) that directly succeeds an explicit memory barrier is a sync op. Such stores are typically used in synchronization schemes like read-copy-update (RCU). Third, any instruction that references a memory-address (such as some memory-allocated variable) that is referenced by another sync op also becomes a sync op. We refer to these operations as unprotected loads and stores, a terminology sometimes used to denote benign data races.

The guarantees we provide by running a replication agent that enforces an equivalent order of sync ops in all replicae are at least as strong as the guarantees that weak determinism provides. To see why, recall that weak determinism enforces a deterministic order of entries into critical sections. Thus, in a correct program, weak determinism will grant mutually exclusive access to related blocks of shared memory in a deterministic order. Though this was historically not the case [45], all modern user-mode programs that run on SMP-systems now implement mutual exclusion using an atomic test-and-set operation [46], [47].

On x86 systems, several instructions provide test-and-set semantics, but they are atomic only if a LOCK prefix is used. Consequently, the first category contains all instructions that may implement the entrance into a critical section as a sync op. Operations that implement the exit from a critical section can either be implemented using atomic test-and-set or exchange operations (second category) or with atomic store operations (first or third category).

On the GNU/Linux platform, all high-level synchronization primitives in the pthreads, OpenMP and C++ standard libraries are based on the mutual exclusion principle. In the pthreads library, e.g., functions such as pthread_mutex_lock and pthread_cond_wait implement mutual exclusion using LOCK CMPXCHG instructions. In the implementation of other high-level synchronization primitives, a variety of atomic test-and-set operations are used. All of them are prepended with LOCK prefixes, however, and all of them are therefore classified as sync ops.

The identification and wrapping of sync ops is currently a partially manual process. First, we disassemble the binary/library to identify explicit memory barrier instructions or instructions with explicit or implicit LOCK prefixes. If no such instructions are present, no further steps are needed. We use debugging symbols to map all of the identified instructions to their originating source line. For memory barriers, we wrap the store that directly succeeds the barrier in calls to our replication API. We identify loads of the same variable and wrap them too. To some extent, we thus fix benign and deliberate data races, such as when developers use such a barrier to set a flag without synchronization.

For source lines that compile into instructions with LOCK prefixes, check whether or not compiler intrinsics are used to express the atomic operations. If not, we insert calls to our replication API before and after the operation. Otherwise, we include an automatically generated header in the source file. This header overrides all known intrinsics that implement atomic operations and inserts the appropriate
replication API calls automatically. It is generated by a simple script that downloads the list of all atomic compiler intrinsics from the GNU GCC website, parses the list and generates the necessary definitions. In addition, we identify other loads and stores of the variables involved in the atomic operations and insert API calls manually.

While this process might seem cumbersome, it is important to note that unless a program or higher-level library implements its own ad hoc synchronization or lock-free algorithms, no memory barriers or instructions with LOCK prefixes will be found. So most programs are supported transparently. For our tests, only four libraries needed modifications. We report extensively on the size of these modifications in Section 3.1.

In many cases, and in particular in portable code, the manual effort to invest in these modifications is very limited. In portable code, compiler intrinsics are used to implement atomic operations, as well as all other accesses to the variables involved in those atomic operations. This is necessary to ensure portability to architectures that do not guarantee the atomicity of aligned loads and stores. To wrap all the necessary compiler intrinsics, it suffices to include our automatically generated header in all source files.

Furthermore, even for programs and libraries that do need more modifications, the patching process can be streamlined to a great extent. C++11 compliant compilers provide a template for atomic operations. With this template, a programmer can mark variables that need to be updated atomically by modifying their type, i.e., by wrapping the type in the std::atomic template. During the compilation, the compiler translates all accesses to such variables such that the appropriate atomic intrinsic is used for every access. Our automatically generated header can then insert the appropriate calls to our replication API automatically by overriding these intrinsics. In summary, if we use a C++11 compliant compiler, the patching effort can be limited to modifying the type of all variables that need to be accessed and updated atomically. In our future work, we plan to extend our current embryonic implementation in LLVM to automate this process completely.

The replication-enabled libraries can replace their original counterparts, in which case they will function correctly and with minimal overhead outside the MVEE context. Alternatively, they can be installed side-by-side with the original ones, in which case the MVEE will intervene transparently in the library loading to load the replication-enabled ones. Our solution hence places a minimal burden on system administrators and users.

2.5 Interaction with the Kernel

Synchronization algorithms often rely on the kernel’s futex API to interact with other threads and processes. The multipurpose synchronization API is exposed through a system call, and is used throughout GNU’s pthreads library for two reasons. First, some functions use the FUTEX_WAIT operation to block the calling thread until the value stored at a specified address changes. In the pthread_mutex_lock function, this operation is used to block the calling thread if the mutex is currently contended. Other functions such as pthread_cond_wait use the FUTEX_WAIT operation to block until an event occurs. Other functions use the FUTEX_WAKE or FUTEX_CMP_REQUEUE operation to signal threads that are waiting for the value at the specified address to change. In the pthread_mutex_unlock function, wake operations are used to wake up threads that are blocked in a related pthread_mutex_lock call. Functions such as pthread_cond_signal or pthread_cond_broadcast use wake operations to signal and wake up threads that are blocked inside a related pthread_cond_wait call.

A potential issue arises when a thread performs a wake operation for which only one other thread should be woken up. If more than one thread is waiting in each replica and the replication agent does not intervene, the kernel might wake up non-corresponding threads in the replica. In a slave replica, the woken thread will then stall indefinitely (i.e., deadlock) at the first atomic op it encounters.

While this issue can be handled in the replication agents embedded in all replica, we chose to tackle the issue from within GHUMVEE’s monitor instead. The monitor allows all replicas to invoke futex calls but it allows only the master replica to actually complete the call. The monitor manipulates the system call number and arguments for the slave replica’s futex calls to have them perform a harmless non-blocking system call instead (such as sys_getpid). The non-blocking call returns immediately from the kernel, at which point the monitor stalls the slave until the master’s futex call has also returned. At that point, the monitor replicates the result of the system call to all slaves and resumes them. This guarantees that the corresponding threads get woken up in all replicas. By implementing the logic in the monitor instead of in the agent, we keep the agent small and fast for its other replication tasks. The additional overhead of going through the monitor is relatively small, given that it already intercepts all system calls invocations and returns anyway.

3 Evaluation

Our implementation of GHUMVEE supports the i386 and AMD64 architectures for the GNU/Linux platform but there are no fundamental restrictions to either the architectures, or the platform: All design options we lifted for GHUMVEE target applications running on top of an unmodified OS running on a commercial off-the-shelf multi-core processor. GHUMVEE has already been tested successfully on desktop programs [24], [21]. In this paper, we focus on evaluating the replication of explicitly parallel benchmarks, on the effort needed to patch libraries and to embed our synchronization replication agent, and on how the replication buffers might enable attacks. As underlying diversity scheme to mitigate code reuse attacks, GHUMVEE implements Disjunct Code Layouts (DCL). DCL ensures that any given virtual address points to a valid executable region in no more than one replica. This policy was demonstrated to provide effective protection against code reuse attacks [24].

3.1 Embedding the replication agent

To evaluate the run-time overhead of GHUMVEE, we used the PARSEC 2.1 benchmark suite. We did not include the facesim and canneal benchmarks because they contain many data races. For canneal, this is hardly surprising as it is based
on data race recovery. We applied minor patches to four benchmarks to eliminate data races or to embed our agent. ferret raced on the cnt_enqueue and input_end variables. Additionally, the imagick library on which ferret depends contained unprotected accesses to the free_segments variable. freqmine raced on the thread_begin_status variable. raytrace used ad hoc synchronization in its AtomicCounter class. vips had an unprotected read and write in the gcosure.c file.

We further applied fixes for bugs which had been reported in the literature or on the PARSEC web site. We configured fluidanimate and streamcluster benchmarks to use the original pthread-based barriers rather than the semantically equivalent but less efficient parsec-based barriers.

On our testing system, the benchmark suite relies on four libraries in which we embedded our replication agent: glibc 2.19, libpthread 2.19, libstdc++ 4.8.2 and libgomp 4.8.2, the default library versions of Ubuntu 14.04. Since libpthread and glibc are built from the same source tree, we treat them as one entity when reporting the required patching effort. For libgomp and libstdc++, we leverage the use of compiler intrinsics as discussed in Section 2.4. Both libraries support specialized targets and more generic targets: libstdc++ supports the so-called i486 and generic CPU targets, while libgomp supports the so-called Linux and POSIX targets. We adapted the makefiles, directory structures, and linker scripts to ensure that the code targeting the generic CPU and POSIX are used instead of the code in support of the more specific i486 and Linux targets, thus ensuring that code relying on compiler intrinsics is used instead of code involving inline assembly. Furthermore, we made sure that the automatically generated header was included in all relevant source files.

For each of the libraries, all of this preparation required editing/executing less than 14 lines of script and source code. The automatically generated header consists of 131 LoC. In addition, in libgomp, 2 lines of code needed to be edited to replace two unprotected load/store operations by atomic ones. So all in all, a very limited effort was required to prepare these libraries: 2*14+2=30 lines of code needed to be edited manually to prepare the two libraries that total about 110k LoC. Moreover, this manual editing was limited to 4 source files out of a total of 673 files.

A considerably larger patching effort was needed to embed our agent in glibc/libpthread, because they use ad hoc synchronization throughout and have many explicit memory barriers and unprotected loads and stores.

Whereas more modern glibc/libpthread ports like the ARM port use compiler intrinsics to implement their atomic operations, the AMD64 and i386 ports do not, presumably because intrinsic support in compilers was not up to par yet when those ports were developed. Instead, the AMD64 and i386 ports rely on inline assembler. With today’s compiler support for intrinsics, the inline assembler can be replaced by intrinsics without performance penalty. In addition, our effort for embedding the replication agent in glibc would have been much reduced in case the inline assembler had already been replaced by the intrinsics. As this is not yet the case, we needed to do the replacement ourselves. For this purpose, we replaced the x86 version of the lowlevellock.h header by the ARM version of that same file. We also deleted the assembly-based x86-specific versions of many pthread functions from the source tree, such that the generic versions of those same functions are used instead.

This did not suffice, however. Contrary to libgomp and libstdc++, glibc’s generic code does not use compiler intrinsics directly. Instead, glibc implements its own series of sync ops, of which some map directly to compiler intrinsics and others do not. So we opted to wrap glibc’s sync ops manually, rather than with the automatically generated header. We manually constructed wrappers span 211 lines of code. We added an additional 175 lines of code to allow ld-linux, which is also built from glibc’s source tree, to still use the original unwrapped macros as needed by GHUMVEE. We therefore needed 386 lines of code in total to wrap all synchronization operations in glibc-libpthread. In addition, we added approx. 261 lines of code to eliminate data races.

Finally, we added 14 lines in one of glibc’s linker scripts to expose our replication agent to other libraries, and added our replication agent itself. The WoC agent, for example, counts no more than 194 lines of C code, while the secure WoC agent counts 167 lines of assembly code and 100 lines of C code.

Excluding the copying and deleting of existing code, as well as our own replication agent, the source code patching effort to prepare glibc/libpthread was limited to 211+175+261+14 = 661 LoC in 60 source code and build files. Compared to the library’s total size of several 100K LoC spread over several thousand source code files, this effort is still fairly limited. And all of it can of course be reused for replicating all applications.

Moreover, since version 2.20, a gradual effort is ongoing in the glibc developer community to replace inline assembler sync ops by their more portable, more generic, more maintainable counterparts in the form of compiler intrinsics. Together with the automated support we are developing as mentioned near the end of Section 2.4, this will reduce the required patching effort significantly in the near future.

3.2 Run-time overhead and scalability
We evaluated our technique on a system with two Intel Xeon E5-2650L processors with 8 physical cores and 20MB cache each. The system has 128GB of main memory and runs the AMDe4 version of the Ubuntu 14.04 OS. For the sake of reproducibility, we disabled hyper-threading and all power saving and dynamic frequency and voltage scaling features. The system runs a Linux 3.13.9 kernel that was compiled with a 1000Hz tick rate to minimize the monitor’s latency in reacting to system calls. We applied a small optional kernel patch (less than 10 LOC) that adds a variant of the sys_sched_yield system call that bypasses GHUMVEE. Other than that, no kernel patches were used. With this small kernel patch, our agents can efficiently yield the CPU whenever they are waiting for preceding sync ops to finish replaying in the slave replicae. This patch improves the performance of our TO and WoC agents in the dedup benchmark but has no significant effects elsewhere.

All benchmarks were compiled at optimization level -O2 using GCC 4.8.2. The native performance of the benchmarks

1. At [http://ghumvee.elis.ugent.be](http://ghumvee.elis.ugent.be) our patches, raw data and scripts are available. GHUMVEE will be open sourced in Q4 2015.
was measured using the original, unpatched libraries that shipped with the OS. GHUMVEE performance was measured using their GHUMVEE-enabled versions.

We measured the execution time overhead of our agents by running each PARSEC benchmark with 1 to 8 worker threads natively as well as in GHUMVEE with 2, 3, and 4 replicate. Using the native PARSEC input sets, i.e., the largest standardized set, we ran each measurement five times, of which we omitted the first to account for I/O and cache warmup. For 1, 2, 4, and 8 worker threads, Figure 5 presents the benchmarks’ execution time as replicated by GHUMVEE, relative to the native versions. For each agent and for each benchmark, Figure 5 shows how the native and the replicated execution (for two replicaes) scales with the number of worker threads.

These figures display several trends. Most importantly, with both WoC agents, many benchmarks (blackscholes, freqmine, raytrace, swaptions, and even streamcluster) can be replicated with little overhead up to 8 worker threads, and in some cases even with 3 or 4 replicate. Other benchmarks (bodytrack, ferret, vips, x264) can be replicated with little overhead up to 4 worker threads. The average overhead of the replication remains below 2x for 2 replicate with the WoC agents, even with 8 worker threads. With more replicate, the overhead clearly increases. This is of course the result of resource contention of the many threads over the limited number of cores. For most of the mentioned benchmarks, it then does not matter too much which agent is used.

For other benchmarks, however, there is a big difference in overhead between the different agents, and there are several benchmarks for which significantly larger overheads and bad scaling are observed.

First, regardless of which agent we use, dedup consistently suffers high performance penalties. The main contributor to this overhead is the high system call density in dedup. When running with 8 worker threads, dedup executes over 123k system calls/second. This density is far greater than in any other program we have tested so far. In the PARSEC suite itself, the highest density we have measured besides dedup was for the vips benchmark (20.9k system calls/second for 8 worker threads). In older benchmark suites such as SPEC CPU2006, the highest density we have measured was around 1k system calls/second for 403.gcc. The high overhead in benchmarks with such high system call densities is unfortunately a fundamental problem of the ptrace API on which GHUMVEE and most other security-oriented MVEEs rely to monitor the behavior of the replicas.

Second, the swaptions and fluidanimate benchmarks, which use fine-grained synchronization, expose a major bottleneck in our PO and TO agents. Both of these benchmarks use a fork/join threading model and frequent, fine-grained synchronization. In both of these benchmarks, all worker threads perform the same tasks and progress at roughly the same pace. While swaptions does not use any explicit synchronization in the application code itself, it does rely heavily on dynamic memory allocation. The dynamic memory allocator in GNU’s libc uses ad hoc synchronization and lock-free algorithms to ensure thread safety. Through libc, swaptions executes more than 398M sync ops when running with 5 worker threads. With 8 worker threads, swaptions performs over 403M sync ops. This corresponds to 4.2M sync ops per second in the native benchmark with 5 worker threads, and up to 7.5M sync ops per second in the native benchmark with 8 worker threads.

In fluidanimate, the situation is even worse. Contrary to swaptions, fluidanimate does invoke our replication agent directly. With 4 worker threads, fluidanimate performs over 1.18B sync ops, which corresponds to over 9.8M sync ops per second in the native benchmark. These sync ops originate mainly from the pthread_mutex_lock and pthread_mutex_unlock functions, which are used to acquire or release one of the 2.31M individual mutexes used in the program. With 8 worker threads, fluidanimate performs over 2.35B sync ops, which corresponds to over 32.9M sync ops per second in the native benchmark. Because the lock and unlock operations are spread over so many different mutexes, there is very little contention in the native benchmark.

In GHUMVEE, however, the TO and PO agents create a lot of contention. Both agents capture the total order of the sync ops in a single circular buffer. To capture this order, the agents acquire a lock before executing the original atomic op and do not release this lock until the operation has been logged in the buffer. These agents therefore effectively serialize the execution of sync ops in the master replica.

A second problem with these two agents is that all replicaes must keep track of their current position in the buffer. This position must be read before the processing of each atomic op, and updated after each atomic op. Every time an update happens on a different core that does not share a cache with the core on which the previous update was executed, the cache line that contains the current position in the circular buffer will be invalidated, and hence cause stalls in the cores’ pipelines.

The combination of these two bottlenecks results in poor scaling for swaptions and fluidanimate. In other benchmarks, the effects of the serialization and the additional cache coherence traffic incurred by our TO and PO agents are less visible. The main reason is that the other benchmarks perform much less sync ops than swaptions and fluidanimate. In the other benchmarks, the highest sync op rates occurred for dedup and vips, with 936K and 644K sync ops per second in the native benchmark resp.

Most importantly, our WoC agents almost completely eliminate the bottlenecks observed in the swaptions and fluidanimate benchmarks. Only in the vips benchmark, these agents cannot avoid a significant serialization overhead when the number of worker threads increases, because it assigns many unrelated mutexes to the same logical clocks. Thus, for this specific benchmark, the PO agent outperforms the WoC agents.

A final trend is that benchmarks that use condition variables do not scale well beyond 6 worker threads. The bodytrack, dedup, ferret, vips and x264 benchmarks all rely on condition variables to signal and to wake up threads. With enough available CPU time, all of the benchmark’s threads can run simultaneously and a thread can be signaled without resorting to sys_futex calls. The five mentioned benchmarks all have heterogeneous threading models and have more than n threads running simultaneously (with n the number of worker threads). For that reason, our machine’s 16 cores simply cannot run all threads in 2 or more replicaes simultaneously.
All in all, the WoC agents perform reasonably well. With 4 worker threads and two replicae, the average slowdown of our MVVEE is only 1.33x with the regular WoC agent and 1.32x with the secured WoC agent. With the TO and PO agents, the average slowdown is 1.73x and 1.64x resp. The high system call overhead in the dedup benchmark is the main contributor to the slowdown. In this configuration, the slowdown in dedup ranges from 2.98x with the WoC agent to 4.19x with the TO agent.

With 8 worker threads and two replicae, the average slowdown is much higher. The slowdown for our TO, PO, WoC and secured WoC agents is 3.69x, 3.42x, 1.99x, and 1.98x resp. For our TO and PO agents, the main cause is the introduced serialization and constant cache invalidations that come with the single circular buffer approach. Our WoC agents eliminate this bottleneck for the most part, but in this configuration, the lack of resources on our test machine becomes a problem. The variations in results for the WoC agents are caused by minor differences in their implementation. The regular WoC agent accesses the synchronization replication buffer directly, whereas the secured WoC agent accesses the buffer indirectly, as we explained in Section 2.3. This indirection slightly increases the cache pressure. As opposed to the regular WoC agent’s C implementation however, the secured WoC agent’s hand-written assembler implementation does not spill any registers to the stack. This optimization slightly reduces the cache pressure. The combination of these two minor implementation differences slightly favors the secured WoC agent in terms of performance.

3.3 Security Evaluation
All of our replication agents rely on a buffer that is shared between all replicae. This buffer is mapped as a read-write
memory segment in the master replica and as a read-only segment in the slave replica. Intuitively, it might seem like a security risk to create such a communication channel between the replica because it can be used to forward information from the master to the slave replica without triggering a monitored RVP. In practice, however, the security risk is minimal.

In principle, it is possible to launch attacks that cause the master replica to write arbitrary data into the buffer, but the master replica cannot instruct the slave replicas to use the arbitrary data in any meaningful way other than to replay synchronization operations, because the data written into the synchronization buffer is only read by the replication agent. We have manually audited our replication agents and verified that they never pass any information they retrieve from the synchronization buffer on to any other part of the program. GHUMVEE further ensures that explicit input, i.e., input retrieved from system calls, is never written into the synchronization buffer.

We do, however, anticipate future MVEE designs in which the MVEE does not arbitrate all system calls that may retrieve input. For example, the recently proposed reliability-oriented VARAN handles system call monitoring and input replication entirely in user space and inside the context and address space of the replica. In such a system, the synchronization buffer could in theory be used as an uncontrolled communication channel, which might aid a compromised master replica in mounting an attack on the slave replicas. Specifically, the compromised master replica could manipulate the return values of its system calls, thereby instructing the slave replica to read further input from the synchronization buffer.

To protect the synchronization replication buffer in this scenario, GHUMVEE forces the buffer to be mapped at different, randomized addresses in each replica. A compromised master replica therefore would not know the exact location of the buffer in the slave replica and it would have to derive the location through information leakage or by guessing. Alternatively, the master replica could try to construct a code reuse attack that invokes the replication agent’s code to read from the synchronization buffer.

GHUMVEE prevents the latter attack with its DCL. The master replica can therefore not mount a code reuse attack: He cannot assume that slave replicas have the same memory layout as the master replica, and if he feeds an address to the replica that points to an executable gadget in the master’s address space, the slave will raise an exception when it tries to execute code at the same address.

Guessing the location of the buffer is hard. GHUMVEE currently use synchronization buffers of $256MiB$, which corresponds to 65536 memory pages. The AMD64 ABI allows user-space applications to use 48 bits for memory addressing but excludes the first memory page, i.e., the page that starts at address 0x0. Therefore, a user-space application may map up to $2^{36} - 2$ memory pages. The chance to blindly guess the location of a $256MiB$ buffer in
one slave replica is therefore \(65536/(2^{36} - 2)\) or \(9.53 \cdot 10^{-7}\).

Forcing the slave replicae to leak the location of their synchronization buffer is not trivial either. While the TO, PO and regular WoC agents do internally store a pointer to the synchronization buffer, GHUMVEE prevents leakage of the pointer from the slaves to the master through the buffer by mapping the buffer as read-only in the slaves. As GHUMVEE intercepts all system calls, it is trivial to prevent a replica from reverting that memory protection.

Leaking the pointer through other channels is still possible, however. We have therefore constructed our secured WoC agent, which significantly reduces the odds of a successful leakage attack. This secured agent does not store a pointer to the synchronization buffer, but instead accesses the synchronization buffer only through an indirection via the \(gs\) segment, as explained in Section 2.3.

A remaining option to consider is the use of a covert channel between the replicae, and to use the MVEE as the medium through which the covert channel communicates. For example, the replicae can deliberately delay each other by exploiting the MVEE’s lock-step execution mechanism. This mechanism dictates that certain operations may only be completed when all replicae attempt to invoke them. The length of the delay can represent information such as individual bits of a pointer value. While it is easy to write programs that intentionally set up and exploit such channels, it is not possible to deploy this technique if the MVEE’s protection policy is properly implemented. In GHUMVEE, DCL prohibits the launch of a code reuse attack (to setup and exploit the covert channels) in the first place.

In conclusion, we believe that our synchronization replication buffers and agents are sufficiently hardened against attacks, even in scenarios where the master replica can forward explicit input to the slave replicae via those buffers.

4 Related Work

4.1 MVEEs

Throughout the last decade, several MVEEs have been presented. Cox et al. first proposed N-Variant Systems, a kernel-space MVEE [16]. Shortly afterwards, Cavallaro presented a proof-of-concept user-space MVEE [18]. Salam et al. then proposed Orchestra, a more advanced user-space MVEE [17]. More recently, Hosek and Cedar presented Mx [20] and VARAN [49], while Maurer and Brumley introduced Tachyon [19]. The latter three systems are not security-oriented MVEEs like the former ones, as they aim at safe testing of experimental software updates, rather than at protecting programs against exploits. The only multi-threaded applications on which VARAN was tested were server applications in which none of the system call behavior depends on the thread synchronization order: Those server benchmark threads perform almost completely independent computations. By contrast, the system call behavior in the PARSEC benchmarks, even our data race free versions, depends very much on the synchronization order.

Without replicating and ordering synchronization events, none of the PARSEC benchmarks can be handled correctly. VARAN’s approach of ordering system calls and signals but not synchronization events, is simply not a generic, reliable solution. While not reported in detail in this paper, we also successfully tested GHUMVEE on all benchmarks on which VARAN and all other mentioned MVEEs were reportedly evaluated. While some of those MVEEs have been evaluated on multi-process applications (such as older version of Apache), GHUMVEE is the first to provide active support for non-deterministic, multi-threaded applications.

4.2 Deterministic Multithreading

Deterministic MultiThreading (DMT) systems impose a deterministic schedule on the execution order of instructions that participate in inter-thread communication, or a deterministic schedule on the order in which the effects of those instructions become visible to other threads. Some DMT systems guarantee determinism only in the absence of data races (weak determinism), while others work even for programs with data races (strong determinism).

Some DMT implementations, especially the older ones, rely on custom hardware [33], [50], [51], [38] or a custom operating system [52], [53]. Of interest to us, however, are the user-space software-based approaches [25], [26], [27], [28], [29], [30], [31], [32], [33], [34], [35].

Software-based DMT systems come in many flavors but essentially, they all establish a deterministic schedule by passing a token. We refer to the literature for an excellent overview of the possible ways to implement the deterministic schedule, as well as their implications [54]. In the remainder of this discussion, we focus on the fundamental reason why DMT systems are incompatible with MVEEs that run diversified replicae: the timing of and prerequisites for the deterministic token passing.

Most DMT systems require that all threads synchronize at a global barrier before they can pass their token. Some of the systems that employ such a global barrier, insert calls to the barrier function only when a thread executes a synchronization operation [25], [26], [27], [28]. This approach is incompatible with parallel programs in which threads deliberately wait in an infinite loop for an asynchronous event such as the delivery of a signal to trigger. Such threads never reach the global barrier. Other DMT systems tackle this issue by inserting barriers at deterministic points in the thread’s execution. These deterministic points are based on the number of executed store instructions [31], the number of issued instructions [35] or the number of executed instructions [34], [33]. All of these numbers are extremely sensitive to small program variations, which makes such systems ill fit for use in diversified replicae.

Conversion [29] does not use a global barrier but, like other DMT systems, it relies on a deterministic token that can only be passed when threads invoke synchronization operations, which again is incompatible with parallel programs in which some threads never invoke synchronization operations. RFDet [32] uses an optimized version of the Kendo algorithm [51] to establish a deterministic synchronization order. Like Kendo however, the order is still based on the amount of executed instructions in each thread, which makes RFDet equally sensitive to program variations.

4.3 Record/Replay

Record/Replay (R+R) systems capture the order of synchronization operations in one execution of a program and then
enforce the same order in a different execution. This can happen offline, by capturing the order in a file to be replayed during a later execution of the same program, or online, by broadcasting the order directly to another running instance of the program. In the absence of data races, R+R systems show many similarities with DMT techniques that impose weak determinism. RecPlay is a prime example of an offline R+R system [39]. During recording, RecPlay logs Lamport timestamps for all pthread-based synchronization operations [40]. During subsequent replay sessions, synchronization operations are forced to wait until all operations with a earlier timestamp have completed. Because it only enforces the order of synchronization operations, RecPlay’s replication mechanism incurs less overhead than preexisting techniques that replicate the thread scheduling order or the order in which interrupts are processed [55]. Moreover, RecPlay assigns the same timestamp to non-conflicting synchronization operations, such that they can also be replayed in parallel.

Loose Synchronization Algorithm (LSA) was one of the first techniques that adopted R+R for use in fault-tolerant systems [56]. LSA designates one of the nodes as the master node. The master node records the order of all pthread-based mutex acquisitions and periodically replicates this order to the slave nodes. These slave nodes then enforce the same acquisition order on a per-mutex basis.

More recently, Lee et al. proposed Respec online replay on multi-processor systems [57]. Oriented towards fault-tolerant execution of identical replicas, Respec purposely records an unprecise order of synchronization operations in the master process and speculatively replays that order in the slave processes. At the end of a replay interval, Respec checks whether the slaves are still synchronized with the master process by comparing their state, incl. their register contents. If not, it rolls them back. While recording, Respec maps synchronization variables onto a statically allocated clock, similarly to our WoC agents. It is doubtful, however, whether Respec’s approach could work in a security-oriented MVEE like ours, in which diversity in the replica makes it hard (if not impossible) to detect whether the replica have diverged at the end of a replay interval.

Other online R+R techniques rely on custom hardware support [38], and hence are not useful for a secure MVEE for off-the-shelf systems.

5 CONCLUSIONS

This paper presented how GHUMVEE was extended to become the first security-oriented MVEE that can replicate parallel programs correctly. We proposed three replication strategies and implemented four replication agents to implement them, one of which does so over a secured communication channel. Our replication agents are conceptually similar to existing tools, but unlike existing tools, they fit within the constraints that a security-oriented MVEE imposes for lock-step monitoring of diversified replicas. Additionally, we proposed a new strategy to embed a replication agent into parallel programs, incl. programs that use ad hoc synchronization primitives, and we evaluated the effort to do so. In the future, we plan to automate this strategy to a large degree.

We extensively evaluated the effect of our MVEE and our replication agents on the PARSEC 2.1 benchmarks on the GNU/Linux platform. With our secure wall-of-clocks agent, the best of the four agents, we achieve an average slowdown of just 1.32x when running the benchmarks with 4 worker threads and 2 replicas.

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