Monitoring Performance Metrics is not Enough to Detect Side-Channel Attacks on Intel SGX

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Abstract

Side-channel vulnerabilities of Intel SGX is driving the research community towards designing low-overhead detection tools. The ones available to date are grounded on the observation that attacks affect the performance of the victim application (in terms of runtime, enclave interruptions, etc.), so they monitor the potential victim and raise an alarm if the witnessed performance is anomalous.

We show that tools monitoring the performance of an enclave to detect side-channel attacks may not be effective. Our core intuition is that these tools are geared towards an adversary that interferes with the victim’s execution in order to extract the most number of secret bits (e.g., the entire secret) in one or few runs. They cannot, however, detect an adversary that leaks smaller portions of the secret—as small as a single bit—at each execution of the victim. In particular, by minimizing the information leaked at each run, the impact of the attack on the application’s performance is significantly lessened—so that the detection tool notices no attack. By repeating the attack multiple times, and each time leaking a different part of the secret, the adversary can recover the whole secret and remain undetected.

Based on this intuition, we adapt attacks leveraging page-tables and L3 cache so to bypass available detection mechanisms. We show how an attacker can leak the secret key used in an enclave running various cryptographic routines of libgcrypt. Beyond cryptographic software, we also show how to leak predictions of enclaves running decision-tree routines of OpenCV.

1 Introduction

Intel Software Guard Extensions (SGX) enables applications to execute in isolation from other software on the same platform, including the OS. SGX-enabled processors run applications in so-called enclaves and provide them with encrypted runtime memory, encrypted storage, and mechanisms to issue authenticated statements on the enclave software configuration. Thus, Intel SGX is particularly suited for cloud deployments because it allows to outsource applications to the cloud, with the assurance that the outsourced application runs unaltered and its data is not available to any (privileged) software on the same host.

Previous work has shown that Intel SGX has a number side-channels that, coupled with an adversary that controls the OS, allow for effective leakage of enclave secrets [9,19,23,30,32]. Alongside attacks, the research community has proposed a number of prevention [5, 8, 12, 14, 25] and detection mechanisms [13, 21, 24]—the former having usually much higher overhead compared to the latter. To the best of our knowledge, all detection mechanisms are grounded on the observation that side-channel attacks affect the performance of the victim application (e.g., by increasing the number of enclave interruptions) and, therefore, signal an attack when the witnessed performance is anomalous.

In this paper, we show that such detection tools may not be effective at detecting side-channel attacks on SGX enclaves. We note that existing detection mechanisms are geared towards an adversary that interferes with the victim’s execution in order to extract the most number of secret bits (e.g., the entire secret) in one or few runs. Such an attack strategy has a significant impact on the victim’s performance, effectively allowing detection mechanisms to notice the anomaly and signal the attack. Our core intuition is to leak smaller portions of the secret—as small as a single bit—at each execution of the victim, so as to minimize the impact on its performance and, therefore, remain undetected.

More specifically, we show that an adversary can profile a victim enclave, thereby identifying the precise moment during the victim’s execution when a specific part of the secret can be leaked via a side-channel attack. For example, if the victim runs the popular square-and-multiply algorithm, we show that the attacker can infer the moment when the $i$-th loop is being executed—i.e., when the $i$-th secret bit is being processed—and run a side-channel attack at that time to leak the secret bit, while generating almost no relevant events for the detection mechanism. By running the victim multiple times and leaking...
a different part of the secret at a time, our technique can recover the whole secret of a victim while remaining undetected. Based on this intuition, we adapt known attacks leveraging page-tables, L3 cache, and a combination of the two, and show their performance on routines of libgcrypt (namely, mpi_powm and mpi_ec_dup_point) used by popular cryptographic primitives such as ElGamal, RSA, and EdDSA. We also apply our attack strategy on non-cryptographic software and show how to leak predictions of enclaves running decision-tree routines of OpenCV [3]. Our results show that our strategy recovers up to 100% of a secret key used in libgcrypt routines, depending on the type of side-channel exploited, and with marginal impact on the victim’s performance (as low as one extra AEX or roughly 40 cache misses per run). In case of a victim using the decision-tree routines of OpenCV to predict handwritten digits of the MNIST data-set [2], our attack strategy can correctly leak around 55% of the predictions (whereas a “standard” side-channel attack, that is easily detected by available tools, reaches 64% of leaked predictions).

We show that an adversary using our attack strategy cannot be detected with existing detection tools such as T-SGX [24], unless one can tolerate a large number of false positives. For example, T-SGX can detect attacks on mpi_powm that leverage page-tables, but at the expense of generating false alarms with probability 0.98. Further, we provide evidence that any detection tool that monitors the performance of the victim is equally likely to fail. We do so by assuming an “ideal” tool (i.e., one with zero performance penalty) that monitors all of the performance metrics proposed in literature and show that even such a tool may not be able to distinguish a benign execution from one where the victim is under attack. Our results, therefore, provide strong evidence that observing the performance of an application to detect side-channel attacks on SGX enclaves may not be feasible, and that effective detection mechanisms are yet to be designed.

## 2 Background & Related Work

In this section, we briefly overview Intel SGX, including side-channel attacks and available defenses.

### 2.1 Intel Software Guard eXtension (SGX)

Intel SGX is a Trusted Execution Environment available in most recent Intel CPUs. It provides an abstraction called enclave that isolates applications and their data from privileged software running on the same host.

Enclave isolation leverages dedicated memory and the hardware performs access control during address translation. However, SGX enclaves share CPU caches with other enclaves and non-enclave processes. Specifically, L1 and L2 caches used by an enclave are shared with processes running on the same core, whereas L3 cache is shared by all program running on the same CPU. Faults and interrupts are handled by the OS via so-called Asynchronous Enclave eXit (AEX). During an AEX, the execution context of the enclave (e.g., stack and CPU registers) is saved to the so-called State Save Area (SSA) within the enclave memory and control is returned to the OS for interrupt handling.

### 2.2 Side-channel attacks on SGX

SGX was shown to be vulnerable to cache-based Prime-and-Probe attacks, whereas attacks that exploit shared memory between the victim and the adversary (e.g., Flush+Reload [33] or Flush+Flush [15]) are not applicable—simply because enclaves have dedicated memory.

Brasser et al., [9] show an adapted version of the Prime-and-Probe attack on enclaves where the victim runs uninterruptedly on the same core of the adversary—by using hyper-threading—and cache monitoring uses hardware performance counters to reduce noise in the observations. Schwarz et al., [23] shows an L3 Prime-and-Probe attack that can recover the private RSA key of an enclave by using only 11 traces and leveraging post-processing.

Previous work has also shown that SGX is vulnerable to other side-channels based on, e.g., DRAM [30] or interrupt logic [29].

CacheZoom [19] shows that a malicious OS can actually increase the effectiveness of Prime-and-Probe attacks by running the victim on a dedicated core—to improve spatial resolution—and by interrupting it frequently (via the Advanced Programmable Interrupt Controller)—to improve temporal resolution.

Since the page-table of an enclave is stored in non-protected memory, a privileged attacker can use it as a side-channel. Previous work [25, 32] leaks the memory access pattern—and in turn secret enclave data—by setting the RESERVED bit of all memory pages so to capture each of the enclave page accesses when page-faults occurs.

**Previous “Stealthy” Side-Channel Attacks.** Previous work proposes side-channel attacks on enclaves that do not cause page-faults—thereby achieving stealthiness despite detection-tools that monitor page-faults. Jo Van et al., [11] monitor the ACCESS bit of the page-table to get the page access sequence of the victim without page-faults. As the ACCESS bit of a page-table is set only the first time the page is accessed (that is subsequent accesses do not modify the bit), the authors of [11] force a TLB shootdown—by interrupting the enclave via inter-process-interrupts—to reset the ACCESS bit. The authors acknowledge that the number of interruptions during their attack is substantially higher than what is to be expected under benign circumstances, and suggest that a detection tool may notice the attack by monitoring enclave interruptions rather than page-faults. Differently, the attack strategy we develop in this paper causes only a few interruptions of the victim and remains undetected.
Detections attacks on L1/L2 page-faults

| Tool         | Performance Overhead | L1/L2 | L3 | page-faults | PTE Monitoring | Our-PF | Our-PFCa | Our-Ca |
|--------------|----------------------|-------|----|-------------|----------------|--------|----------|--------|
| Varys [21]   | ~15%                 | ✓     |    | ✓           | ✓              | ✓      | ✓        | ✓      |
| Déjà Vu [13] | ~4%                  | ✓     | ✓  | ✓           | ✓              | ✓      | ✓        | ✓      |
| T-SGX [24]   | ~108%                | ✓     | ✓  | ✓           | ✓              | ✓      | ✓        | ✓      |
| “Ideal” Tool | 0%                   | ✓     | ✓  | ✓           | ✓              | ✓      | ✓        | ✓      |

Table 1: Detection tools for side-channel attacks on Intel SGX. A “✓” (resp. “×”) denotes the fact that the tool can (resp. cannot) detect the attack.

Wang et al., [30] show that enclave interruptions can be minimized if TLB shootdown is achieved by using a sibling hypervisor that probes memory addresses whose TLB entries are conflicted with the ones of the victim enclave. While the attack developed by [30] cannot be detected by monitoring enclave interruptions, it requires the adversary and the victim to run on the same core. As such, the attack is not viable in case of detection tools that enforce core-reservation like Varys [21] or Déjà Vu [13]. Differently, our attack strategy does not require the adversary to run an hypervisor on the core where the victim is running.

### 2.3 Defenses

Existing defenses can be categorized either as prevention or detection techniques.

Prevention techniques remove leakage by, e.g., preventing untrusted threads from using the victim’s cache [12], rewriting programs to remove secret-dependent page accesses [25] or secret-dependent memory accesses [5], pre-fetching memory [14], or shuffling memory [8]. These approaches usually incur high overhead [5, 14], and sometimes can only prevent specific types of side-channels [12].

Differently, detection techniques have usually lower overhead. To the best of our knowledge, all detection techniques proposed in literature (see Table 1) monitor the execution of the victim application and signal an attack in case of deviations from a “normal” execution. Varys [21] prevents L1/L2 cache-based attacks with core-reservation; at the same time, Varys detects attacks based on page-faults or interrupts by monitoring the number of AEXs so that an alarm is raised if their frequency is too high. We note that Varys is currently part of a commercial product named SCONE and its source-code is not available. Déjà Vu [13] detects attacks based on page-faults or interrupts by monitoring the execution time of the enclave—based on the assumption that AEXs due to faults or interrupts, increase the execution time of a program. Déjà Vu instruments the basic blocks of the enclave code to measure their execution time and an attack is detected if the total time deviates from the one of an execution in a benign environment. An incomplete version of Déjà Vu is available on github1; we made contact with the authors to obtain the missing code, but they are no longer maintaining the project. T-SGX [24] makes use of Transactional Synchronization eXtensions (TSX) to detect page-faults. When an interrupt or fault is thrown within a TSX transaction, TSX aborts and executes a user-defined handler. The handler of T-SGX keeps tracks of the number of AEXs and raises an alarm if they reach a given threshold. The source code of T-SGX is available on github2.

In Table 1, we also include an “ideal” tool that incurs no performance penalty for the application that uses it, and that monitors all of the performance metrics proposed in literature. We will later show (in Section 5) that even if such a tool were available, it cannot detect an adversary using our attack strategy.

### 3 Stealthy Side-channel Attacks

#### 3.1 System Model

We assume that the adversary has the victim code available (e.g. the code belongs to a library or an open-source implementation), controls the OS where the enclave is running, and can execute the victim enclave arbitrarily many times.

Such assumptions are similar to the ones found in related work on controlled-channel attacks [19, 30, 32] and capture a realistic cloud scenario where an application is uploaded by its owner to the cloud provider, and part of the application code (e.g., a decryption routine) runs in an enclave. After attestation and secret provisioning by the application owner, the cloud provider can (re-)start the application or trigger the routine running in the enclave arbitrarily many times.

#### 3.2 Main Intuition

As mentioned earlier, detection tools for (known) side-channel attacks leverage the fact that attacks are likely to alter the performance of the victim application. Thus, available detection tools monitor the performance of the potential victim, and signal an attack if the witnessed performance is anomalous.

Our intuition to bypass these tools, is to minimize the effect on the victim’s performance by “spreading” the attack across multiple runs. In particular, we show that an adversary can extract specific portions of the secret, as small as a single bit, at each run of the victim enclave. By minimizing the information leakage at each run, the impact of the attack on the victim’s performance is also lessened—so that the

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1https://github.com/schuan/dejavu

2https://github.com/sslab-gatech/t-sgx
Victim Thread

| Segment 1 | Segment 2 | Segment 3 | Segment n-1 | Segment n |
|-----------|-----------|-----------|-------------|-----------|

Attack Thread

1st run

2nd run

... 

nth run

Figure 1: Spreading the attack across multiple runs. Different portions of the secret are leaked at each run.

detection tool notices no anomaly. This strategy is repeated for a number of times—each time leaking a different portion of the secret—to eventually recover the full secret.

In particular, let the enclave secret be \( s = s_1, \ldots, s_n \), where each \( s_i \) could be a single bit or multiple ones. Also, assume the victim code to be split into \( n \) segments \( S_1, \ldots, S_n \), such that segment \( S_i \) processes \( s_i \). As shown in Figure 1, the application is executed \( n \) times. During the \( i \)-th run, the attack thread launches a side-channel attack while the victim is executing segment \( S_i \), in order to leak \( s_i \). As the attack only runs for a small time-window, the victim’s performance is only marginally affected.

Developing the aforementioned strategy requires the adversary to mount a side-channel attack only during the time-window when the victim is executing code segment \( S_i \). One option would be to precisely control the victim’s execution by using single-stepping frameworks like SGX-Step [28]. Nevertheless, side-stepping the victim generates a large number of page-faults, so a tool that monitors the number of AEXs could easily spot the attack. To remedy this, we take a different approach and design an offline profiling phase to learn the time-interval when the victim is executing a specific code segment \( S_i \). In the following, we detail the offline profiling phase and the design choices we made to minimize errors.

3.3 Application Profiling

Let \( T_i \) denote the time when the victim starts the execution of code segment \( S_i \). Note that a segment is a logical execution unit and different segments may execute the same code, but on different portions of the secret. For example, in the square-and-multiply routine each segment corresponds to one execution of the main loop and processes a single bit of the secret exponent.

In an ideal scenario, the execution time of each code segment is constant, i.e., \( T_{i+1} - T_i = c \). Thus, segment \( S_j \) starts at time\( T_j = (i - 1) \cdot c \), for some constant \( c \). More generally, the execution time of a code segment may depend on the code itself, as well as the portion of the secret it processes. Thus, we model the execution time of segment \( S_i \) as a function \( t_i(s_i) \), and set the start time of segment \( S_i \) as \( T_i = \sum_{j=0}^{i-1} t_j(s_j) \).

As an example, Figure 2 shows a simple code segment with a conditional branch on the \( i \)-th bit of variable \( secret \) and three different function calls (\( m \), \( g \), and \( k \)). In case there are no conditional branches nor loops in functions \( m \), \( g \), and \( k \), we can use constants to model their execution time as \( c_m, c_g \) and \( c_k \), respectively. Thus, \( t_i(s_i) = c_m + s_i \cdot c_g + (1 - s_i) \cdot c_k \). In case any of the functions \( m \), \( g \), \( k \) has a loop or a conditional branch, we would recursively profile its execution time in a similar fashion.

Once we have the function \( t_i(s_i) \) that models the execution of \( S_i \), we assess its values by running \( S_i \) multiple times, and by using a different assignment of \( s_i \) each time. For example, to measure the execution time of the code in Figure 2, we run the segment twice: once with \( s_i = 0 \) and once with \( s_i = 1 \). Note that the enumeration of all possible configurations of the part of the secret processed by a code segment is feasible because each segment is likely to process only one or a few secret bits.

Measuring Execution Time. Since SGX cannot execute privileged instructions (e.g., rdstc), we cannot directly inject time measurement instructions into code segments. Furthermore, when measuring the execution time of each code segment we cannot interrupt the enclave, as context switches between enclave and non-enclave code incurs extra overhead compared to context switches between regular processes [7, 31]. We, therefore, create a logical clock by means of a timer thread. We inject instructions at the start and end of each segment, to set a binary variable at a memory address Addr outside of the enclave memory. A separate timer thread continually gets the system timestamp using rdstc and checks the value of the variable at Addr. If the variable is set to 1, the timer thread remembers the current timestamp and reset Addr to 0. By measuring the time interval between two reads of Addr that returned 1, we can infer the time required to run one code segment.

Stabilizing Execution Time. The execution time of arbitrary code on a general-purpose machine is far from deterministic due to the complexity of the host and other software running on it. As in previous work [30], we reduce the noise due to other software on the same host by reserving a core for

```
void compute_on_s(char[] p, unsigned int secret)
{
    int tmp = m(p);
    if ((secret & (1 << i)) != 0) {
        g(tmp);
    } else {
        k(tmp);
    }
}
```

Figure 2: Sample code segment.

3.3 Application Profiling

Let \( T_i \) denote the time when the victim starts the execution of code segment \( S_i \). Note that a segment is a logical execution unit and different segments may execute the same code, but on different portions of the secret. For example, in the square-and-multiply routine each segment corresponds to one execution of the main loop and processes a single bit of the secret exponent.

In an ideal scenario, the execution time of each code segment is constant, i.e., \( T_{i+1} - T_i = c \). Thus, segment \( S_j \) starts at time\( T_j = (i - 1) \cdot c \), for some constant \( c \). More generally, the execution time of a code segment may depend on the code itself, as well as the portion of the secret it processes. Thus,
Although techniques such as speculative execution may still improve attack accuracy.

Improving Attack Accuracy. The accuracy of our technique relies on the correct estimation of \( T_i \)—when we start the side-channel attack to learn \( s_i \)—and the correct guess of \( s_j \).

Clearly, an error when estimating \( T_i \) leads to a misalignment between the attack and the victim that, in turn, leads to unpredictable errors in inferring the secret \( s_i \). An error when inferring \( s_i \) may lead to an error in the estimation of \( T_j \) for \( j > i \) since the start time of segment \( S_j \) may depend on the value of \( s_i, ..., s_{j-1} \).

One possible option (among others) to avoid misalignments between the victim and the attack is to rely on page-faults. In particular, the attacker may invalidate a page that is required by the victim at the start of \( S_i \), so that a page-fault is thrown when the victim starts executing that segment. Alternatively, alignment errors may be corrected by attacking multiple consecutive segments at a time by using a sliding window. This basic idea is shown in Figure 3. Let \( w_i \) be the window attacking segments \( S_{i-w_i+1}, ..., S_i \) so to obtain bits \( s_{i-w_i+1}, ..., s_i \) and, without loss of generality, assume the step of the window to be 1. Then, we compare the guess for bits \( s_{i-w_i+1}, ..., s_{i-1} \) obtained when attacking window \( w_i \), with the guess for the same bits obtained when attacking window \( w_{i-1} \) (i.e., when attacking segments \( S_{i-w_{i-1}}, ..., S_{i-1} \)). If the two bit sequences match, then we assume that window \( w_i \) is well aligned and treat the guess for the last bit of the window (i.e., \( s_i \)) as valid; otherwise we assume \( w_i \) is not aligned with the victim and discard the guess for \( s_i \).

Note that attacking larger windows may have an impact on the victim’s performance that could allow a detection mechanism to spot the attack. Also, our attack with \( w = n \) becomes similar to a “standard” side-channel attack that tries to leak all secret bits at once.

In order to improve the accuracy of our technique, we can also increase the number of times we attack a given segment. That is, we run the victim \( k \) times and run the attack on the same segment \( S_i \) (or segment window \( w_i \)). We therefore obtain several samples for \( s_i \) and use heuristics to improve the accuracy of our guess.

Figure 4: \texttt{mpi\_powm} used in ElGamal, RSA and DSA.

4 Leaking secrets of libgcrypt

We now show how to instantiate the strategy described earlier on cryptographic routines \texttt{mpi\_powm} and \texttt{mpi\_ec\_mul\_point} of libgcrypt. The former is the modular exponentiation routine used in a number of cryptographic primitives, including ElGamal, RSA, and DSA; \texttt{mpi\_ec\_mul\_point} is the elliptic curve point multiplication routine used by EdDSA. For both victims, we leverage a side-channel based on time as shown in [17, 30], one based on the memory access pattern as shown in [11, 25], and a combination of the two. In the following, we use libgcrypt version 1.7.0. Nevertheless, the side-channels we exploit are present in \texttt{mpi\_powm} up to version 1.8.6, and in \texttt{mpi\_ec\_mul\_point} up to version 1.7.5.
4.1 Side-channels of \texttt{mpi\_powm}

Figure 4 shows the code of \texttt{mpi\_powm}. The routine has two side-channels, one based on time and another based on memory access pattern.

The loop (line 8 \textasciitilde{} 29) consumes one bit of the secret exponent per iteration and executes an extra computation (line 11 \textasciitilde{} 28) if that bit is 1 (line 10). Thus, an adversary can infer the secret bit of the exponent being processed, by inferring the time to complete one loop iteration. Note that if \texttt{as\texttt{ec}} is 1, then the exponent is stored in secure memory, and the conditional branch is always executed to eliminate side-channels. However, if \texttt{xvalue} is provided as input by the user (e.g., when the key-pair is generated from a passphrase), then \texttt{libgcrypt} does not store the exponent in secure memory so that side-channels are not eliminated.

Alternatively, the secret bit of the exponent can be leaked by monitoring access to memory pages that store the code required by the 1f-branch of the routine. Let A, B, C be the addresses of \texttt{mpi\_powm}, \texttt{mpi\_mpih\_sqrt\_n\_basecase} and \texttt{mpihelp\_mul}, respectively. One iteration of the loop where the exponent bit is 1, shows a memory access sequence like ABCAC|AB, whereas if the exponent bit is 0, the observed memory access sequence is like ABC|AB. In these examples, memory accesses after 1 belong to the next iteration of the loop. Also, note that \texttt{mpi\_mpih\_sqrt\_n\_basecase} calls \texttt{mpihelp\_mul}, so there will always be an access to address C after B. One could infer the memory access sequence either by observing page-faults or cache accesses.

4.2 Profiling of \texttt{mpi\_powm}

In order to profile \texttt{mpi\_powm}, we define each iteration of the main loop as one segment. Let \( s_i \) be the \( i \)-th exponent bit consumed in segment \( S_i \). One iteration of the loop in \texttt{mpi\_powm} computes on \( x_p \), \( r_p \) and \( s_i \). We found no branches nor loops in \texttt{mpih\_sqrt\_n\_basecase}, so its runtime is constant and we denote it as \( c_{\text{base}} \). Thus, runtime of \( S_i \) with \( s_i = 0 \) is simply \( c_{\text{base}} \). If \( s_i = 1 \), the code executed (lines 11 \textasciitilde{} 28) has two branches. The first one is a conditional branch that, depending on the value of \texttt{bsize}, may run either \texttt{mpihelp\_mul} or \texttt{mpihelp\_mul\_karatsuba\_case}. We found that both paths take the same time so we model this time as a constant \( c_{\text{mpihelp}} \). The second branch depends on \texttt{xsize} and \texttt{msize}. However, we found that the time taken to run \texttt{mpihelp\_divrem} is negligible, so we just ignore it. In a nutshell, the time to run segment \( S_i \) is \( t_i(s) = c_{\text{base}} + s_i \cdot c_{\text{mpihelp}} \) and \( T_i = (i-1) \cdot c_{\text{base}} + \sum_{j<i} s_j \cdot c_{\text{mpihelp}} \).

4.3 Page-faults

We start by describing an instantiation of our attack strategy that only uses page-faults and that leverages the timing side-channel of the victim; we denote this attack variant as Our-PF.

4.4 Page-fault and Cache

We now describe another attack variant that leverages both page-faults and cache misses, that we refer to as Our-PFCa. As Our-PF Ca only leverages a side-channel based on memory access pattern, it could be used on routines that have no side-channel based on time.

Similar to Our-PF, we use one page-fault to stop the enclave at the beginning of each segment. Then, we use a
Prime-and-Probe attack on L3 cache to infer the secret bit processed during the execution of that segment. We use L3 since most detection tools prevent L1/L2 attacks by ensuring that no untrusted thread runs on the same core of the thread being monitored (see Table 1).

The workflow of Our-PF Ca is shown in Figure 6. Let \( c_0 \) and \( c_1 \) (with \( c_0 < c_1 \)) be the time it takes to run one loop of mpi_powm with secret bit 0 and 1, respectively. We stop the enclave at \( T_i - c \) by making the page of mpi_powm_basecase unavailable at that time; next, we resume the victim and wait for \( r \) clock ticks to make sure that computation on mpi_powm_basecase is over. Now, the goal is to measure whether the next call to mpi_powm_basecase happens after time \( c_0 - r \) or \( c_1 - r \). To do so, we start a Prime-and-Probe attack on the address of mpi_powm_basecase, for a period of \( c_0 - r \). We construct the eviction set of the Prime-and-Probe using techniques from previous research [18]. Figure 7 shows the time to access the target cache set when the secret bit is 1 (a) or 0 (b). Here, it is clear that the victim has accessed the cache line of mpi_powm_basecase if the access time of the attacker to the eviction set is larger than 1000 ticks. The first peak in each figure denotes the start of the \( i \)-th iteration, while the shaded area denotes the interval of \( c_0 - r \) ticks during which we run the Prime-and-Probe attack. Note that if \( s_i = 1 \) (Fig. 7a) we do not witness any access to mpi_powm_basecase while running the Prime-and-Probe attack. In case \( s_i = 0 \) (Fig. 7b) we witness access to mpi_powm_basecase as the routine moves to the next iteration of the loop. Once we learn \( s_i \), we compute \( T_{i+1} \) accordingly, and move on to attack the next segment.

### 4.5 Cache-only

The attack strategies above use page-faults to temporally align the victim and attack threads. We now show how to run cache-only attacks on the victim enclave. In the sequel, we refer to this strategy as Our-Ca.

Note that, using only cache to leak a specific portion of the victim’s secret may be difficult because the adversary thread may not be aligned with the one of the victim; nevertheless, Our-Ca is particularly effective with detection tools that monitor enclave exits (AEXs) [22, 24] as it enables the leakage of the secret without interrupting the victim at all.

Our-Ca works by starting a Prime-and-Probe attack on the address of mpi_powm_basecase right before \( T_i \) and for \( c_1 \) ticks—the number ticks required to complete the loop iteration when the secret bit is 1. If the attack thread experiences a peak in the time to access the target cache set, followed by a sufficient number of lows, we conclude that \( s_i = 1 \), whereas if the attack thread experiences two close peaks, we conclude that \( s_i = 0 \).

A considerable challenge when using Our-Ca lies in the fact that small errors when estimating \( T_i \) leads to unpredictable cache patterns. This is shown in Figure 8. In Figure 8(a), the attack starts at the right time and the witnessed cache pattern does indeed support a correct guess of \( s_i \). Differently, in Figure 8(b) the attack starts late and
void _gcry_mpi_ec_mull_point (mpi_point_t result, gcry_mpi_t scalar, mpi_point_t point, mpi_ec_t ctx)
{
    /* Other operations */
    if (mpi_is_secure (scalar))
    {
        /* Oblivious Implementation */
        for (j=nbits-1; j >= 0; j--)
        {
            _gcry_mpi_add_points (result, result, point, ctx);
            if (mpi_test_bit (scalar, j))
            {
                _gcry_mpi_add_points (result, result, point, ctx);
            }
        }
    }
    else {
        for (j=nbits-1; j >= 0; j--)
        {
            _gcry_mpi_dup_point (result, result, ctx);
            _gcry_mpi_add_points (result, result, point, ctx);
        }
    }
    /* Other operations */
}

Figure 9: mpi_ec_mul_point used in EdDSA.

the adversary (mistakenly) estimates $s_i$ to be 0. Finally, in Figure 8(c), the attack starts early, preventing the adversary from estimating the value of the secret bit.

We correct alignment errors between adversary and victim by using a sliding window technique as explained in Section 3. That is, when attacking window $w_i$ (i.e., segments $S_i$ to $S_i$) we start the Prime-and-Probe attack right before $T_{i-1}$ and we run it for $wc_i$ ticks—i.e. until the end of segment $S_i$. Next, we consider the estimate of $s_i$ as valid only if the estimate of $s_{i-1}$...$s_{i-1}$ matches the estimate of the same bits when attacking window $w_{i-1}$. Finally, we also repeat the attack on the same window a number $k \geq 1$ of times in order to obtain multiple guesses for the same bit and use an heuristic to infer its actual value.

4.6 Attacks on mpi_ec_mul_point

We now briefly discuss how to adapt our attack strategy to the mpi_ec_mul_point routine of libgcrypt used in EdDSA. For each signature, this subroutine is used to compute scalar multiplication with a nonce that, if leaked, allows the recovery of the signing key. Note that our attack extracts one secret bit for each execution of the victim; hence, if the victim picks a fresh nonce at each execution, two bits extracted by our attack would be completely uncorrelated. Nevertheless, EdDSA is deterministic [27] and the nonce is computed as function of the message to be signed and the signing key. Hence, by feeding a fixed message to the signing routine we ensure that the nonce is always the same and can extract one of its bits at each execution.

Figure 9 shows the code of mpi_ec_mul_point. Note that the same routine is used to process both the nonce and the signing key (referred to as scalar in both cases). The leakage-free code (line 7 ~ 10) is used when processing the signing key, whereas the else-branch is taken to process the nonce. In the latter case, a secret-dependent branch (line 13) can be abused to leak one bit of the (secret) nonce. Once the nonce and the corresponding signature are available, the signing key can be computed.

Profiling mpi_ec_mul_point. Let segment $S_i$ be the i-th iteration of the loop. We found that there are no conditional loops nor branches in gcry_mpi_ec_dup_point so we model its execution time with constant $c_{base}$. In case the bit of scalar being processed is 1, the routine calls another constant-time function called mpi_ec_add_point and we model its execution time with constant $c_{add}$. Therefore, the running time of the i-th loop iteration is $T_i = c_{base} + s_i \cdot c_{add}$, and the start time of the i-th segment is $T_i = (i-1) \cdot c_{base} + \sum_{j=1}^{i-1} s_j \cdot c_{add}$.

In practice, we must also accommodate for the first iteration of the loop that takes the if-branch; this iteration must fetch mpi_ec_add_point and its callees from the main memory and incurs in a time increase that we model with $c_{miss}$, Thus the start time for the i-th segment becomes $T_i = (i-1) \cdot c_{base} + \sum_{j=1}^{i-1} s_j \cdot c_{add} + (\sum_{j=1}^{i-1} s_j \cdot c_{miss})$. This extra time for the first loop that processes a 1 bit does not show in mpi_pomw, as the secret-dependent call to mpi_sqrt_n_basecase is also called in other function before mpi_pomw.

When attacking mpi_ec_mul_point we use the page containing mpi_ec_dup_point to stop the enclave at the beginning of the target segment. Further, we target mpi_ec_dup_point when launching the Prime-and-Probe attack.

5 Results for libgcrypt

We instantiated the relevant routines of libgcrypt 1.7.0 in SGX, by using Panoply [26] to delegate system calls out of enclave. Our experiments were carried out with Ubuntu 18.04 on an Intel E3-1280 with 4 physical cores and 32GB RAM.

To assess the effectiveness of our adapted attacks despite existing detection tools, we compiled and run the two cryptographic routines using T-SGX with some engineering efforts. The abort handler in its code simply recomputes after aborts, and we modified the handler to forward page-fault handling to the system.

As the other two detection tools available in literature are not open-source—Varys is part of a commercial product and Déjà Vu is no longer maintained—we also evaluate the effectiveness of our strategy on enclaves equipped with an “ideal” tool. We assume the latter to require no code instrumentation (hence, “ideal” in terms of performance overhead), to monitor all of the performance metrics proposed in literature (i.e., number of AEX, cache misses, and execution time), and to raise an alarm if the witnessed performance is anomalous. We note that cache misses are typically monitored via performance counters—a feature that is not currently available for SGX enclaves. Nevertheless, previous work has shown that (non-SGX) applications could use cache-misses to detect cache-based attacks [10]; hence, we include the
number of cache misses among the performance metrics that are monitored by the ideal tool to emulate the possibility that it becomes available to future SGX applications.

5.1 Profiling Accuracy

We start by measuring the execution time of one loop of the victim routines—recall that a loop of \texttt{mpi\_powm} and \texttt{mpi\_ec\_mul\_point} is a code segment as defined in Section 3. We do so by running each loop 100 times with secret bit 0 and another 100 times with secret bit 1.

Our results show that, in case of using the ideal tool, a “0-loop” of \texttt{mpi\_powm} takes on average 46.4k clock ticks ($\sigma = 493.6$), while a “1-loop” takes on average 92.9k clock ticks ($\sigma = 122.2$). When instrumented with T-SGX, \texttt{mpi\_powm} takes slightly longer: 48.3k clock ticks ($\sigma = 530.1$) for a 0-loop and 103.2k clock ticks ($\sigma = 251.3$) for a 1-loop.

Routine \texttt{mpi\_ec\_mul\_point} with the ideal tool, takes on average 15.4k clock ticks ($\sigma = 378.1$) for a 0-loop, and 39.2k clock ticks ($\sigma = 284.9$) for a 1-loop. When instrumented with T-SGX, \texttt{mpi\_ec\_mul\_point} nearly double its computation time: it takes 38.8k clock ticks ($\sigma = 631.3$) and 92.0k clock ticks ($\sigma = 376.1$) for 0-loop and 1-loop, respectively.

Once we have the running times for 0-loop and 1-loop iterations, we validate the accuracy of our profiling technique by checking whether we can stop the enclave at the start of each loop. To do so, we fix a random 256 bit secret and we execute the enclave 256 times, each time stopping it (by leveraging a page-fault) at time $T_i - c$ (with $i = 1, \ldots, 256$ and $c = 5,000$ clock ticks). In order to learn the ground truth, we inject a counter into the code to keep track of the number of loop iterations thus far. This experiment is repeated 20 times and we report the results in Table 2.

Our evaluation shows that stopping at a specific code segment an enclave running \texttt{mpi\_powm} with the ideal tool is more accurate (93.17%) than achieving the same if the enclave runs the routine instrumented with T-SGX (85.08%). This is because we may lose synchrony with the victim as T-SGX restarts a transaction. We observe the same behavior for \texttt{mpi\_ec\_mul\_point}: 81.74% for the version using the ideal tool and 68.19% for the version instrumented with T-SGX. A comparison between \texttt{mpi\_powm} and \texttt{mpi\_ec\_mul\_point} shows lower accuracy for the latter. This is because one loop of \texttt{mpi\_ec\_mul\_point} takes less time to complete compared to a loop of \texttt{mpi\_powm}—therefore, it is harder to hit the start of a specific loop iteration.

5.2 Attack Accuracy

We evaluate the accuracy of the three adapted attack variants in recovering secret bits. To do so, we fix a random 256 bit secret and, for $i = 1, \ldots, 256$, we recover the secret bit $s_i$ by attacking the corresponding code segment 9 times (i.e., for 9 times we run the enclave and launch the side-channel attack from $T_i$ until $T_{i+1}$). Given the 9 samples, we determine the secret bit based on majority voting. We repeat the experiment 10 times and show the average accuracy and standard deviation in the column “Attack Accuracy” in Table 3 and Table 4 for \texttt{mpi\_powm} and \texttt{mpi\_ec\_mul\_point}, respectively.

Table 2: Accuracy when stopping the victim enclave at the beginning of a specific code segment.

| Function                  | Accuracy                  |
|---------------------------|---------------------------|
| \texttt{mpi\_powm} (w/ ideal tool) | 93.17% ($\sigma = 5.49\%$) |
| \texttt{mpi\_powm} (w/ T-SGX)              | 85.08% ($\sigma = 8.90\%$) |
| \texttt{mpi\_ec\_mul\_point} (w/ ideal tool) | 81.74% ($\sigma = 4.52\%$) |
| \texttt{mpi\_ec\_mul\_point} (w/ T-SGX)              | 68.19% ($\sigma = 1.40\%$) |

For comparison purposes, we also report the accuracy of “standard” side-channel attacks using either page-faults [30] or L3 cache [17,30]. A standard attack refers to an attack that runs throughout the whole execution of the victim in order to recover the largest number of secrets bits in one execution. In case of standard attacks we also repeat the attack 9 times and use majority voting to decide the value of each secret bit. Note that, in case of routines instrumented with T-SGX, a standard page-faults attack does not work as the detection tools raises an alarm shortly after the attack starts.

Table 3 and Table 4 show that our attack strategy can recover between 60% and 100% of the enclave secret, depending on (i) the type of side-channel exploited, (ii) the detection tool used by the victim (either T-SGX or an ideal tool), and (iii) the number of consecutive code segments attacked per run. For example, we can extract around 80% of the secret using Our-PF and up to 100% of the secret using Our-Ca. Our experiments also show that attack accuracy decreases when the victim is instrumented with T-SGX: this is likely due to the noise introduced by T-SGX when restarting transactions that abort before completion.

Comparing the accuracy of attacks on \texttt{mpi\_powm} with the accuracy when attacking \texttt{mpi\_ec\_mul\_point}, we note that Our-PF performs better on \texttt{mpi\_powm} and this is because the time difference between a 1-loop and a 0-loop in that routine is sharper than the time difference of the loops in \texttt{mpi\_ec\_mul\_point}. Nevertheless, attack variants that use cache are more accurate on \texttt{mpi\_ec\_mul\_point} as the cache side-channel is more noisy when attacking \texttt{mpi\_powm}. Furthermore, cache-only attack with larger windows (e.g., \(w = 9\)) provide very good results.

We also assess the impact of the number of samples $k$ we obtain for each secret bit on the attack accuracy. As shown in Figure 10, increasing $k$ improves accuracy that, however, plateaus around $k = 9$ for most of the attack variants.

Finally, we assess the impact on accuracy when relying on a sliding window to reduce alignment errors in cache-only attacks. Recall that attacking a single segment at a time by
only using cache side-channels may lead to poor results due to the difficulty of aligning the victim and attack threads (see Section 3). In our experiments, a cache-only attack on one segment at a time resulted in an average accuracy over 20 runs of 46.64% ($\sigma = 3.84\%$) for mpi_powm with the ideal tool; the accuracy for the version instrumented with T-SGX was 48.44% ($\sigma = 3.4\%$). The same experiment when attacking mpi_ec_mul_point showed an average accuracy of 51.64% ($\sigma = 1.98\%$); the accuracy for the version instrumented with T-SGX was 50.3% ($\sigma = 3.7\%$). By using the sliding window technique described in Section 3, we improve accuracy as shown in Figure 11. In particular, a window of size $w = 9$ allows to recover the full secret when attacking mpi_powm, whereas the same result can be achieved with a window of size $w = 5$ for mpi_ec_mul_point. This is because, the cache side-channel is less noisy in mpi_ec_mul_point, as explained before.

### 5.3 Effectiveness Against Detection Tools

We now assess the effectiveness of T-SGX and the ideal tool described above in detecting our attack strategy. Recall that T-SGX monitors the number of AEXs whereas the ideal tool monitors all of the performance metrics proposed in literature, namely number of AEXs, execution time, and cache misses.

We collect the aforementioned performance metrics, when the victim is under attack, as well as when the victim is running in a benign environment either (i) alone or (ii) while another process is running on the same machine. For the latter, we used either Redis—a key-value store—and we mimic a realistic workload as in [4], or GCC while building a large project. Finally, we record the required performance metrics while attacking the victim with standard side-channel attacks.

On the one hand, reported figures on routines instrumented with T-SGX provide us with evidence of the effectiveness...
of our adapted attacks when the victim is equipped with existing tools. On the other hand, results of the experiments with an ideal tool allow us to reason about effectiveness of our adapted attacks with respect to any tool that monitors the performance of the potential victim. For each scenario, we run the victim 1,000 times and report the average and standard deviation in Table 3 and Table 4.

For each of the considered scenarios and for different values of detection thresholds, we also measure recall (ratio of true positives over all positive cases) and specificity (ratio of true negatives over all negative cases). The recall metric is used to measure security: a perfect tool should have recall equal to one, and any smaller value means that the tool is not detecting some attacks. Specificity measures usability: a perfect tool should have specificity equal to one, and any smaller value means that the tool is raising some false alarms. In a real deployment, detection thresholds should be set so that both recall and specificity are as close as possible to one. In the following, we show that, for both T-SGX and the ideal tool, high specificity and high recall cannot be achieved at the same time—which shows that such tools cannot detect an adversary that use our attack strategy.

**Detection with T-SGX.** In this set of experiments, we run the victim instrumented with T-SGX along with GCC to mimic a realistic multi-threaded workload. We vary the threshold number of AEXs before T-SGX raises an alarm, and measure specificity and recall for each threshold.

Figure 12a shows results for mpi_powm. No threshold in the range we consider ($2 \leq t \leq 10$) provides specificity greater than 0.03—hence, one should set the threshold to a value much larger than 10 to avoid false alarms. At the same time, all of the attack variants exhibit zero recall for $t \geq 5$. Figure 12b shows very similar results for mpi_ec_mul_point. Specificity does not increase above 0.03 for $t \leq 10$ and all of the attack variants reach zero recall for $t \geq 6$.

**Detection with the ideal tool.** In case of the ideal tool, we consider detection based on both the number of AEXs and the number of cache misses. For each scenario, we run the victim either along with GCC—to mimic a multi-threaded workload—or along with Redis—as an exemplary application to mimic a memory-intensive workload. Further, we vary the detection threshold—either the one of number of AEXs or the one of number of cache misses—and measure specificity and recall for each threshold.

Figure 13a and Figure 13b show results when the tool is monitoring AEXs and the victims are mpi_powm or mpi_ec_mul_point, respectively. No threshold in the range we consider ($2 \leq t \leq 10$) provides specificity greater than 0.37 in case of mpi_powm and 0.55 in case of mpi_ec_mul_point. At the same time, all of the attack variants go undetected (recall is 0) if $t \geq 5$ for both victims.

Figure 13c and Figure 13d depict our results when the tool is monitoring cache misses and the victims are mpi_powm or mpi_ec_mul_point, respectively. For both victims, specificity is 0 for thresholds up to 800; therefore, the tool raises a false alarm at every run and one should set much higher thresholds to avoid false alarms. However, all of the attacks go undetected if $t \geq 800$ for mpi_powm or $t \geq 600$ for mpi_ec_mul_point.
Specificity (No Attack with GCC)

Detected by monitoring execution time. A detection tool may monitor the execution time to decide whether the application is under attack. This is for example the case of Déjà Vu [13]. Results from Table 3 and Table 4 show that standard page-fault attacks almost double the execution time and would be likely detected by tools such as Déjà Vu. Differently, our attacks cause minimal increase of the victim’s execution time (below 2%); as such, it is challenging for Déjà Vu or similar tools to detect them.5

6 The Case of OpenCV

We now adapt our attack strategy to known side-channels of decision-tree routines [20] in order to attack the decision-tree routine of OpenCV [3], a well-known computer vision library. Similar to previous work [1], we use the MNIST [2] data-set and assume an application consisting of an enclaved execution of OpenCV’s decision-trees to detect handwritten digits.

Figure 14 shows the code of OpenCV’s decision-tree traversal function DTreesImpl::predictTrees. This function traverses the tree and, depending on the input image, accesses different nodes (nodes[nidx]), resulting in different page accesses. We use page-faults to infer the pattern of page accesses and leak the prediction output. To capture the access pattern of different input images, we rely on an offline analysis of the routine and observe memory page access patterns of different nodes[nidx] for prediction output. These memory pages are set as fault during runtime, to infer the prediction output. To reduce the number of page-faults during attacks, we eliminate memory page accesses that cannot identify two unique secrets. For example, if prediction “1” and “2” produces access sequence ACD and ACE, respectively, then we use only page A, D and E for inferring the prediction.

We ported OpenCV’s decision-tree library (the i.e., core and ml modules) to SGX by forwarding system calls (e.g., fread and fopen) outside enclave. For training the decision-tree, we use 60,000 samples from the MNIST data set. During the inference phase, the trained decision-tree model is first loaded into the enclave and then used to recognize 100 input images at a time from a set of 10,000 test images. The execution time of image recognition is almost independent of the input image, as the tree is almost balanced. Therefore, we can model the execution time as $T_i = T_i + c$, where $c$ is a constant value. In our experiments, we found out that $c$ is roughly 9.7k clock ticks ($\sigma=975.3$).

Profiling and Attack Accuracy. In our experiments, we were able to stop the enclave at the time of the $i$-th invocation of DTreesImpl::predictTrees around 8 out of 10 times (84.8% ($\sigma=7.85\%$)). The corresponding attack accuracy is reported in Table 5—our attack that only leverages page-faults (Our-PF) is only slightly less accurate than a standard page-fault attack.

Effectiveness Against Detection Tools. To analyze the effectiveness of our strategy against detection tools, we measure specificity and recall for different AEX thresholds. We only assume the victim is equipped with the ideal detection tool—as T-SGX supports only C, we could not instrument OpenCV using T-SGX. Figure 15 shows that no threshold value can achieve high specificity and high recall at the same time. In particular, if the detection threshold is smaller than 17, then specificity falls below 0.84 (i.e., a false alarm is raised 2 out of 10 times). At the same time, a detection threshold equal to or bigger than 11 allows attacks to go undetected (recall=0.003). We also note that Our-PF causes no noticeable overhead in terms of cache misses or execution time. We conclude that an ideal detection tool—one that monitors number of AEXs, cache misses or execution time—may not be able to tell an attack that uses our strategy from a benign run of the victim enclave.

5Note that for a 256 bit secret, attacking a single segment takes roughly 6 ms (see for example the running time of mpi_powm in Table 3). In this case, recovering the full secret requires approximately 1536 ms. In case the adversary attacks each segment $k$ times to obtain higher accuracy, the total attack time increases to roughly 13.8 seconds.
7 Concluding Remarks

Our findings show that an adversary can bypass existing tools that monitor performance metrics to detect side-channel attacks on SGX enclaves, by exfiltrating small portions of a secret at each run of the victim. This is particularly relevant for existing cloud-based deployments where enclaves can be restated multiple times by the cloud operator/administrator.

One possible countermeasure would be to tune existing detection tools in order to spot an adversary that uses our tools that monitor performance metrics to detect side-channel attacks on SGX enclaves, by exfiltrating small portions of a secret at each run of the victim. This is particularly relevant for existing cloud-based deployments where enclaves can be restated multiple times by the cloud operator/administrator. Another countermeasure could be to prevent arbitrary restarts of the victim enclave. Namely, the victim enclave may be allowed to execute only upon receiving an authenticated request from, e.g., an authorized user. This is a workable option only if the enclave is used by few users (e.g., in a scenario where a user outsources expensive computations to his enclave deployed in the cloud). Nevertheless, this option is not workable when the enclave provides a “public” service. For example, if the enclave hosts a TLS server [6] or a password-hardening service [16], it is challenging to differentiate between an authorized request from a honest user and another issued by the adversary acting as a honest user.

Intel SGX, however, does not provide freshness of state information sealed to disk. A malicious OS can, therefore, bypass such a tool by providing stale state to the enclave.
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