Master research Internship

Bibliographic report

Software Fault Isolation using the CompCert compiler

Domaine: Cryptography and Security

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Abstract: Software Fault Isolation (SFI) is a software method which allows one to isolate the execution of possibly dangerous modules, hence preventing any risks of corruption of our system due to these modules. Even though SFI has multiple benefits it still has difficulties dealing with Returned Oriented Programing attacks which aim at diverting the control flow through the return addresses. During the internship we implemented a flexible solution against these attacks with the certified compiler CompCert. Our approach transforms the stack layout in order to easily know the locations of the return addresses enabling us to protect them from being overwritten illegally. We made an implementation that was able to successfully protect our system from buffer overflows. In return we have a diminution of the efficiency depending on the quantity of protection added. However multiple ideas exist to reduce this overhead like having a finer detection of the vulnerabilities of the compiled code with the use of static analysis.

Contents

1 Introduction

IT projects regularly need to use external modules in their programs. However these modules can come from unknown sources and we don't have any guarantees that they don't contain any bugs or malicious code. Software Fault Isolation is an approach which purpose is to face this issue by verifying certain security properties during the execution of the dangerous modules.

A rising interest for the techniques of Software Fault Isolation (SFI) has appeared in the field of cyber security. SFI is a mean to protect the memory of a program with software solutions. The idea is that a protected program can load in its own memory space dangerous modules without them being able to compromise its execution. To make it happen, SFI isolates the dangerous modules in reserved areas of the memory called sandboxes. This operation is realised by a code generator that modifies the code of the dangerous modules in order to have them comply with the security properties defined by SFI. Afterwards the modified code of the modules goes through a verifier that checks that the modifications introduced before are present and valid.

SFI techniques are applicable in multiple fields. An implementation has already been made for Google Chrome [9][6] called NativeClient. We can also imagine using SFI for critical systems that have to execute external modules. For example computer clusters or system kernels could require the use of SFI to safely execute external modules.

SFI has multiple strong points which explains its current appeal. Firstly, we can note high performances due to the execution of the modules natively in the address space of the protected program. This trait is quite important since the most popular memory protection mechanisms use hardware solutions and suffer from great overheads. Another benefit of SFI is that it requires a relatively small *Trusted Computing Base* (TCB). Indeed only the verifier needs to be in the TCB and not the code generator. Having a small TCB means that only a minor portion of the code (the verifier) needs to be trusted to have a correct implementation of SFI.

However SFI has still difficulties to deal with the protection of `ret` instructions that use return addresses stored on the stack. A big part of modern attacks, like *Return Oriented Programming* attacks, focus on deviating the control flow through these return addresses. Therefore it is absolutely necessary to have a good protection method for return addresses. The SFI solutions found in the literature are not easily adaptable or bring a lot of overhead to the compiled modules.
Hence we propose a solution which is more flexible and faster to address this specific issue of SFI. The idea is to modify the stack structure to have an easy way to figure out the return addresses locations in the stack. We set the frames size to a constant value in order to have the return addresses separated by a fix offset. Afterwards we modify the dangerous code by adding runtime checks which take advantages of the new stack layout to prevent any instruction to overwrite the return addresses.

This internship report will firstly present Software Fault Isolation, its principles, the existing implementations and especially a version using CompCert the certified compiler. CompCert is also be the compiler we chose to implement our solution in order to defend programs against ROP attacks. We will then present an overview of our approach with fixed frames size with a discussion of our idea. Afterwards we will detail the implementation we made of our approach with CompCert during the span of the internship. To conclude we will evaluate the different results obtained with our implementation and suggest ideas to improve our solution in the future.

2 Software Fault Isolation

We introduce here *Software Fault Isolation* (SFI) which inspired us the idea to protect return addresses through fixed stack frame size. SFI aims at protecting a main program from the different modules that it will need to use. These modules will be loaded in the same memory space as the main program but in a confined area called *sandbox*. The SFI mechanism is composed of two elements: a code generator and a verifier. The generator transforms the assembly code of the hazardous modules so that they will be constrained in the sandbox. The verifier operates just before loading the modules in the memory. It checks if the SFI transformations introduced by the generator are still present and valid. For the rest of the document we will reserve the word "program" to refer to the code protected by SFI and "module" to refer to the hazardous code.

2.1 Principle

The main principle behind SFI was first presented in the work of Wahbe and al. [8]. Later works that will be introduced Chapter 2.2 are all based on the foundations of SFI detailed here. The implementation described here was realised for a RISC architecture like MIPS or *Alpha*.

SFI considers that a malicious code is effectively contained in the sandbox if these three security properties are true:

- **Verified code**, only instructions that have been checked by the verifier will be executed
- **Memory safety**, malicious modules will not do any *write* or *jump* operations out of the sandbox
- **Flow control integrity**, every flow control transfer from hazardous modules to the main program is identified and verified

The first property protects us against self-modifying code which can bypass the SFI measures. *Memory safety* prevents any illegal access to the memory of the protected program. The last property allows us to authorize only licit interactions between the program and its modules. SFI forbids any call from malicious modules that can modify the flow control of the program. If the flow control is compromised, it can lead to an unexpected behaviour of the program which we want to avoid.

The code generator transforms the assembly code of the hazardous modules so that they respect the security properties presented before. The generator is integrated to the compiler which will create a sandboxed executable. Afterwards this executable will be checked by the verifier before being loaded into the memory. The verifier checks that the transformations introduced by the generator are present and valid. If the verification fails the module will be rejected and will not be executed. We can note that we only need to trust the verifier to avoid running any dangerous module. It is one advantage of SFI, only the verifier needs to be in the *Trusted Computing base* (TCB).

2.1.1 Code generator

To protect a program from its modules, the generator will restrain every write and jump instructions of the modules to addresses of their sandbox. The generator has to face three issues to do so. The first one is to introduce protection mechanisms before every dangerous instructions.

For example preventing any destination address of jump instructions to be located out of the sandbox. Secondly, we have to make sure that these protection mechanisms can't be avoided. Finally, the transformations injected have to authorized only legal calls from the sandbox to the protected program by using entry points specified by the latter. For example, Google Chrome only allows its modules to use a specific interface to interact with the browser. This way the modules can't disrupt the flow control of Google Chrome easily.

Confining memory accesses. The main program memory should avoid being corrupted by its modules. SFI aims at isolating these modules in a reserved area of the program memory called sandbox. The sandbox is a contiguous memory area which size is a power of two. These requirements ease the confinement of the modules in their sandbox by allowing the use of bit arithmetic which accelerates the process. In fact two sandboxex are allocated, the first one for the code segment and the second for the data segment. In these conditions we only need to verify that the most significant bits of the targeted address match those of the sandbox. For example, if we allocate to a sandbox the memory `[0xda000000 - 0xdaffffff]`, then all the addresses which most significant bits match `0xda` are located in the sandbox. Thus all the SFI transformations will restrain the memory writes and jumps to the addresses of the sandbox area. In our example, we will limit ourselves to the addresses which most significant bits are `0xda`, this sequence of bits is also called **tag**. Each tag is specific to an unique sandbox and this designation will be used repeatedly in our report.

During the code generation phase every direct memory addressing will be easily detected with static analysis. The transformation will be simple, the targeted addresses will simply have their most significant bits replaced by the tag of the sandbox. Another possible reaction is to reject the code if it is explicitly unsafe. The real issue appears when faced with a write or a jump using indirect memory addressing. Indeed in these cases the destination address is stored in a register and we can't access this value during the compilation. However to address this situation SFI injects in the code of the dangerous module runtime checks that are called *sandboxing*.

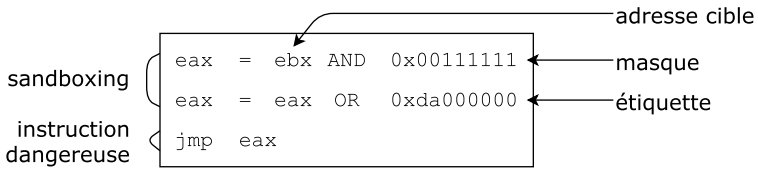


Figure 1: Pseudo code of the sandboxing operation

Figure 1 represents an example of the sandboxing operation. The sandboxing starts with a masking operation which sets the most significant bits of the address stored in the register `ebx` to zero. Afterwards the second instruction writes the tag of the sandbox on the bits it just reinitialised before. Hence we are sure that the `jmp` instruction will target a location in the sandbox of the dangerous module. We can note that the sandboxing does not change the behaviour of the module if the targeted address was already in the sandbox. This principle is called *transparency*, SFI transformations does not change the execution of safe programs. For the write instructions the principle is the same. We inject the sandboxing instructions before every write that can endanger our program.

Protection of sandboxing mechanisms. We made sure in the previous section that any dangerous modules can't either jump or write on a location out of its sandbox. Now we also want to protect the sandboxing operations to prevent any malicious code to bypass the runtime checks inserted by SFI. Using the example Figure 1, we could imagine code which directly jumps on the `jmp eax` instruction. To protect the sandboxing, the solution is to establish dedicated registers exclusively used for sandboxing. These registers won't be available anymore for the rest of the code. Sandboxing requires three dedicated registers for each sandbox. First register is used to keep the mask value (`0x00111111` in Figure 1). Second register is reserved to store the tag of the concerned sandbox (`0xda000000` in Figure 1). And the third dedicated register is used to manage the operations contained in the sandboxing, in our example Figure 1 it would be the register `eax`. This way during the whole execution of the dangerous module `eax` only stores addresses of its sandbox. Then even if malicious code can jump directly to the instruction `jmp eax` we will still be sure that it stays in the sandbox. Worst case scenario would be that the value stored in `eax` was wrong and the dangerous module crash or has unexpected behaviour. As long as the protected program is not affected the protection is considered successful. These dedicated registers are never used by the rest of the code and their values can't change except during sandboxing operations.

Since we have two sandboxes, one for the data and one for the code we then have a total of six dedicated registers. However SFI manages to reduce the number of dedicated registers to five by sharing the same mask for both the sandbox.

We can question ourselves on the efficiency of SFI when we remove five registers for the execution of the normal code. However, in the modern RISC architecture like MIPS or Alpha there are generally 32 registers. Moreover the experiments [8] show that removing five registers for the gcc compiler impacts insignificantly the efficiency of the programs tested.

Controlled interactions with the protected program. It is necessary for SFI to also control the different interactions that dangerous modules have with the main program. Without restrictions, malicious modules could, for example, make function calls with wrong parameters which could compromise the state of the main program. To avoid such situation, SFI needs the main program to define an interface which describes all the authorized entry points available to external modules. This interface also includes the range of authorized values for the different parameters. For example Google Chrome has created an API which can be called by the different modules to access the resources of the browser. SFI then transforms the modules so that every interaction with the main program is done through connectors called *stubs*. These stubs are part of the Trusted Computing Base (TCB) of SFI. In relation with the interface of the main program, the stubs make sure that function calls from dangerous module to the main program are licit. If they are not, the stubs reject the function calls. System calls are controlled the same way, the stubs first check the call and if it is authorized they transmit the system call to the main program. Then the main program executes the system call to the kernel and returns the results to the module.

2.1.2 Code verifier

The verifier is the last element of the SFI chain. Consequently it is necessary for it to be part of the TCB contrary to the code generator. Even if the code generator is flawed the verifier rejects every executable which does not comply with the SFI requirements. Nevertheless the verifier relies on the generator work to do the checking. The verifier analysis just asserts that the SFI transformations made by the generator are present and valid. Therefore it's not possible for a

binary which does not go through the generator to be validated by the verifier even if the binary is safe.

The first step of the verification is to disassemble the given binary. The implementation of Wahbe et al. [8] was made for the architectures MIPS and Alpha. In these architectures all the instructions are 32 bits long. This particularity makes the disassembly easier since we just need to treat sequences of 32 bits. The second step is to check that the SFI transformations are present and valid. The first condition is that the registers dedicated to store the mask and the tags should never be modified. Afterwards we need to check that the sandboxing operations are done correctly. For that we define in the modules some sandboxing areas. These areas begin every time that one of the dedicated registers is modified. Then these sandboxing areas are closed when we do a jump or a write with a dedicated register or when there are no instructions left. To validate a sandboxing area we verify that the dedicated register used for the jump or the write will have an authorized value at then end of the area. In other words we make sure that the dedicated register stores a value which most significant bits are equal to the tag of the sandbox.

2.1.3 Pros and cons

SFI aims at isolating the execution of dangerous modules in a designated area of the protected program memory. For this, SFI is composed of a code generator tasked with producing a sandboxed binary and a verifier which checks if the executable complies with the SFI properties. It is imperative for the verifier to be part of the TCB in order to have SFI only accept safe code. We can note that to be functional SFI does not need an extensive TCB, only the verifier needs to be part of it. Another advantage is that SFI is not language dependant, the approach does not depend on the source language of the code. In return SFI is architecture dependant, the transformations will not be the same on x86 or on MIPS. Another con is that the sandboxed code is also less efficient and also bigger. Nevertheless only the dangerous modules are sandboxed which usually represent a minor part compared to the whole execution. A last inconvenient is that the presented implementation only includes MIPS and Alpha architecture but later works adapt SFI to other architectures.

2.2 SFI for other architectures

In the work of Wahbe et al. [8] the methods used was made for the RISC architecture and especially MIPS and Alpha. We detail in this section the necessary modifications created to adapt SFI to x86 and ARM.

2.2.1 SFI for CISC architectures

Contrary to RISC architecture in a CISC architecture the instructions may have different lengths. This particularity introduces two issues for SFI. First of all it is possible for an instruction to jump in the middle of another instruction which would totally change the flow of the program. Moreover the disassembly phase of the verifier becomes much more complex and the assembly obtained is not always reliable. In RISC architecture there is a natural alignment since all the instructions are 32 bits long. The processor then always makes sure that it jumps on an address respecting the 32 bits alignment. This way the code is protected from this kind of threats. In CISC it becomes a serious issue which can't be ignored by SFI.

To face this problem, Pittsfield [5] proposes to create an artificial alignment using chunks of memory of constant size. To have this approach work the sandboxed code needs to verify some additional properties:

1. The sandbox is divided by chunks of constant length equal to a power of two
2. Every instruction targeted by a jump has to be located at the beginning of a chunk
3. Every function call has to be located at the end of a chunk
4. It is forbidden for an instruction to overflow on multiple chunks
5. Every chunk which are not full are filled with `no-op` instructions

The first rule allows us to create an artificial alignment on the chunks length. The rule number 2 enforces the property that every destination of a jump instruction is aligned. In other words with memory chunks of sixteen bits, every jump has its target address with the four least significant bits equal to zero. The rule 3 guarantees us that every return address from a function call is the beginning of a chunk and then is aligned. The rule number 4 is necessary for the instructions to respect the artificial alignment.
From all these rules we can deduce an atomicity property for the memory chunks. More explicitly, since it is not possible to jump in the middle of a chunk we can be certain that all the instructions of a memory chunk are executed atomically. This atomicity property is a big gain for SFI because it allows us to free ourselves from the use of dedicated registers. Indeed these dedicated registers were used in order to prevent bypassing the sandboxing of dangerous write or jump. By using the artificial alignment we can make sure that it is not possible to avoid the sandboxing operation and execute the dangerous instruction directly. We just need to gather all the sandboxing operations with their dangerous instruction in the same memory chunk, the atomicity of the chunks then protects our mechanism.

The downside of this approach is that the artificial alignment slows down the execution of the modules concerned. Benchmarks carried out show a decrease of the average performance by 21% compared to a non-modified module.

2.2.2 NativeClient the implementation of SFI by Google

Google developed NativeClient [9] which is included by default on its browser Google Chrome. Chrome uses SFI to execute possibly dangerous modules. These modules are executed natively and show better performances than with JavaScript. NativeClient is certainly the most complete implementation of SFI nowadays.

NativeClient for x86-32. The first version of NaCl (Native Client) was developed for the architecture x86-32 [9]. This implementation combines the usage of x86 memory segments with ideas of the previous approach for CISC architecture. In order to make NaCl complies with the security properties of SFI, Google created multiple rules presented Figure 2.

NaCl uses x86 memory segments to confine the data memory area in continuous memory area. These memory segments are then used as a natural sandbox for the writing instructions. Therefore we don't need to do sandboxing operations for write instructions since it is handled natively by the processor.

C1	Once loaded into the memory, the binary is not writable, enforced by OS-level protection mechanisms during execution.
C2	The binary is statically linked at a start address of zero, with the first byte of text at 64K.
C3	All indirect control transfers use a `nacljmp` pseudo-instruction (defined below).
C4	The binary is padded up to the nearest page with at least one `hlt` instruction (0xf4).
C5	The binary contains no instructions or pseudo-instructions overlapping a 32-byte boundary.
C6	All *valid* instruction addresses are reachable by a fall-through disassembly that starts at the load (base) address.
C7	All direct control transfers target valid instructions.

Figure 2: NaCl rules (Bennet Yee et al. [9], Table 1)

Afterwards we still need to confine the jump instructions to the sandbox. NaCl uses the memory segment `CS` which is a code segment. We force the segment to start at the address 0 like indicated by the rule C2 of Figure 2. This way the most significant bits of the addresses in the sandbox are equal to 0 and we can then reduce the sandboxing to two operations instead of three like presented Figure 3. Moreover NaCl guarantees us that the destination address is a multiple of 32 (rule C5 of Figure 3). This technique is inspired from the approach using memory chunks seen earlier. This operation is called `nacljmp` by NaCl (rule C3 of Figure 3).

```
and         %eax, 0xffffffe0
jmp         *%eax
```

Figure 3: `nacljmp` the reduced sandboxing operation (Bennet Yee et al. [9])

We can see that the rules C1 and C6 of Figure 3 assure us that the disassembly of the binary is reliable. C1 in particular allows NaCl to protect programs against auto modifying code. If auto modifying modules were accepted they would render SFI pointless. This way with a reliable disassembly NaCl can make sure that only the analysed instructions can be later executed. This core property to SFI is rarely evoked in the other works and NaCl is the only one which implements it explicitly.

NaCl also chooses to forbid system calls, instructions that modify the memory segments state and replace the `ret` instruction with a `pop` and a `nacljmp` to the return addresses. This last measure was taken to control the inherent vulnerability of `ret` instruction. `ret` uses an address stored in the stack as its destination address. Since the return addresses are read from the stack they are subject to concurrency from the different threads. Therefore even if we add a runtime check on the values of return addresses we do not have the guarantee that the value has not changed between the sandboxing and the execution time of `ret` (Time Of Check Time Of Use).

NativeClient for X86-64 and ARM. NaCl aims to be used with the popular browser Google Chrome, so Google needs to adapt its product to the most used architecture which are x86-64 and ARM (smartphones). These two versions of NaCl are evolutions from the x86-32 version

modified to fit x86-64 and ARM specificities. For example, these two architectures do not have memory segments and NaCl was forced to reuse more classical methods of sandboxing for the write instructions. These implementations do not bring many novelties to SFI, hence we will not detail the work done but more information can be found on their publications [6].

To conclude on NaCl, the two implementations showed a slight decrease in efficiency of 5% for ARM and 7% for x86-64. These satisfying results prove that SFI is viable for modern architectures to execute dangerous modules without risks. As we said before, NaCl is included in Google Chrome and it is possible to test it by downloading in the web store of Google compatible applications like Quake which is a classic of the video games world.

2.3 SFI using a certified compiler

Until now we only saw implementations of SFI based on the traditional model: a code generator followed by a verifier belonging to the TCB. Another approach we explain here is the usage of a certified compiler. These compilers have the particularity to keep the semantics of the source code. Therefore it is possible to do the SFI transformations in a higher level than the assembly and have the guarantee that the executable produced will keep the properties introduced by the SFI transformations made at a high level. We will specifically talk about the certified compiler CompCert [4] and its usage for SFI.

2.3.1 CompCert the verified compiler

CompCert. CompCert is a certified compiler which supports almost all the C defined by ISO C 99. This compiler can produce code for multiple architectures like PowerPC, ARM and x86-32. Its particularity is that it was formally proven and written with Coq the proof assistant. These proofs allow us to trust the compiler and that the semantic of the compiled programs will be preserved during the compilation.

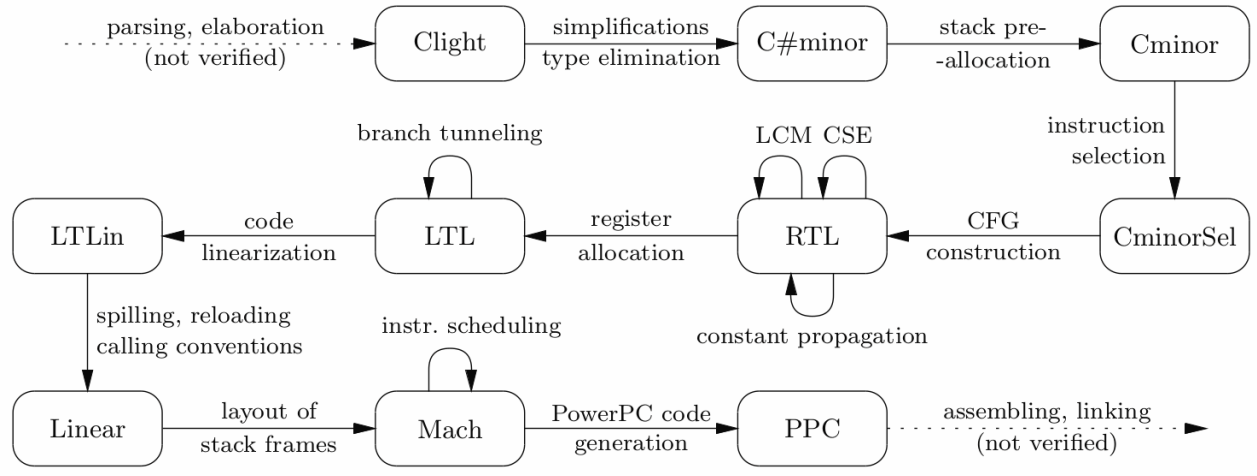


Figure 4: Compilation routine of CompCert (Xavier Leroy [4], Figure 1)

CompCert has a standard compilation routine which is presented Figure 4. It has got some classic optimizations like constant propagation or elimination of common subexpressions. Perfor-

mances of code produced by CompCert are correct, on PowerPC it reaches 90% of the efficiency of `gcc` with optimizations at level 1.

CompCert verifies a theorem of correctness which certifies that every program compilable by CompCert will keep its semantic. More explicitly, every safe source program S (*safe* here means that all its code is well defined in ISO C 99) will be compiled in an assembly code C which verifies the same properties than S. In the case of a source program with undefined behaviour (in the semantic of CompCert C) we will have no guarantee on the correctness of the produced code. For example an operation without semantic in CompCert is an access out of the bounds of an array.

Memory model of CompCert. CompCert uses an abstract representation of the memory for its semantics [1]. This model is used to prove the correctness of the multiple compilation steps of CompCert. Therefore all the intermediary languages of CompCert share this memory model.

The memory is represented by blocks of finite size and each are given a block number b. These blocks are separated and to access a precise location in each block we use an offset δ. Pointers are then represented by the couple $(b,\ \delta)$. It is also possible to represent function pointers which are represented by pointers with negative block numbers b. It can be surprising that such abstract model of the memory can give us enough finesse to represent our SFI transformations. However the fact that memory blocks are disjointed is enough to isolate a part of the memory. Thus one of these memory blocks will be reserved as the sandbox of our implementation.

2.3.2 SFI with CompCert

We present an implementation of SFI with CompCert that was done by Joshua A. Kroll, Gordon Stewart et Andrew W. Appel [3]. The approach is composed of two parts. Firstly, all the SFI transformations are done in the intermediary language Cminor. We will then get a sandboxed Cminor code verifying the properties of SFI. The correctness theorem of CompCert guarantees that if the sandboxed Cminor is *safe* in the definition of CompCert then the produced code will also verify the properties of SFI. This way, the verifier used in the traditional approach is not necessary anymore since CompCert gives us equal guarantees. Hence this approach of SFI will not need a verifier in its TCB which is an advantage.

Cminor. Cminor is an intermediary language created for CompCert. It is similar to a simplified version of C where we remove the type of the variables. Cminor was chosen for the SFI transformation because it is the lowest level language in the compilation steps of CompCert which is still architecture independent. This way, SFI transformations done in Cminor will be architecture independent as well. Moreover Cminor is also low level enough to be an intermediary language for source languages other than C. For these different reasons Cminor is the ideal candidate for the sandboxing operations of SFI.

Specification of the SFI transformation. SFI techniques give us the following requirement. Any implementation of SFI should be able to transform any dangerous module into a module which respects the properties imposed by SFI. Therefore even for a semantically ill-defined module, CompCert-SFI is supposed to produce SFI compliant code. In this case the sandboxed Cminor needs to verify these two properties:

- the sandboxed Cminor confines all the memory accesses to its sandbox

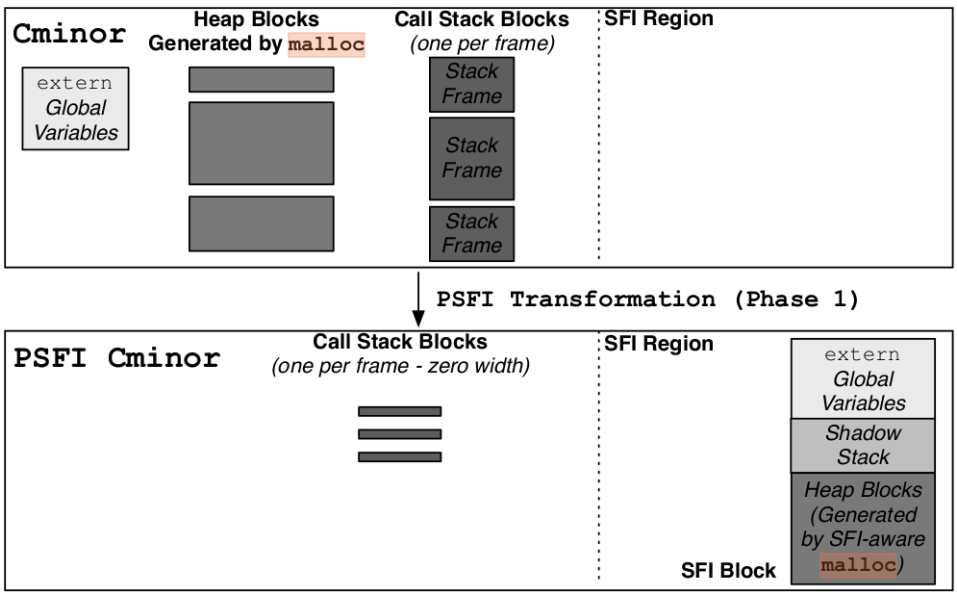


Figure 5: *Sandboxing* of a module in CompCert (Joshua A. Kroll et al., Figure 1)

- the sandboxed Cminor needs to be safe semantically in order to be able to use CompCert correctness theorem

This means that even if the source program is semantically ill-defined for CompCert, the SFI transformations need to make it semantically safe even if it modifies its behaviour. However if the source program is already secure the SFI implementation will not modify its behaviour.

Masking in CompCert. CompCert-SFI uses the memory model of CompCert to prove that the SFI properties injected in Cminor still hold after the compilation. To represent the masking operation in the semantic of CompCert a native CompCert *mask* function was created. This *mask* function was then axiomatized to be able to prove that the SFI transformations does not change the semantic of secure programs. For example one of the axiom of the *mask* function is "If a pointer already points to the memory block of the sandbox then *mask* does not change its value".

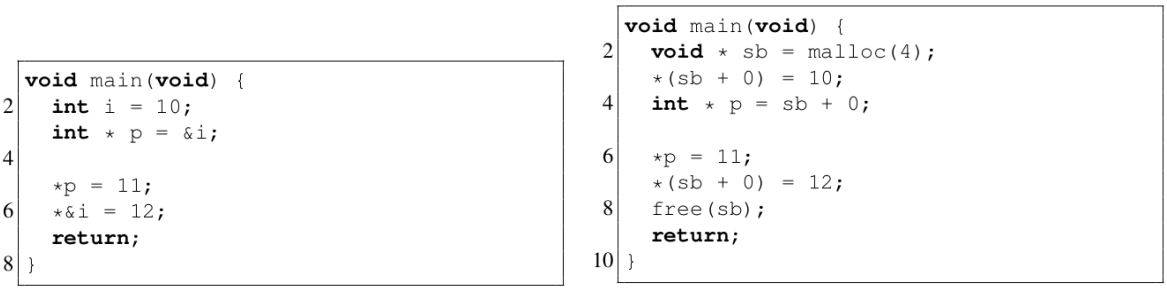
SFI transformations There are two ways in CompCert Cminor to allocate memory. The first one is to use an external function like `malloc`, the other way is to allocate frames in the stack. Then we need to ensure that the Cminor dangerous code can only allocate memory in the sandbox like in Figure 5. Therefore CompCert-SFI transforms the dangerous Cminor with multiple steps:

- *Creation of a secondary stack*, this second stack is created in the SFI area. Hence all the local variables of the dangerous modules are stored in this second stack and stay in their sandbox. To create such stack in the sandbox, we create a modified `malloc` function that only allocates memory in the sandbox. This SFI-`malloc` is used to allocate the different frames of the secondary stack like in Figure 6.

- *Replacement of* `malloc`, like we said earlier an SFI-`malloc` was created that only allocates memory in the sandboxes. This SFI transformation replaces all the classic `malloc` used with

this SFI-`malloc` to confine the modules to their sandbox.

- *Masking the memory accesses*, we inject the *mask* function described earlier before all possibly dangerous instructions.
- *Masking the function call*, memory areas are specially reserved to identify and control function pointers that could compromise the control flow of the program. Each memory area is dedicated to one authorized function of the protected program. Therefore the function pointers are redirected to these memory areas which control the interactions between the modules and the protected program.
- *Production of semantically safe code*, we want that all the code produced by SFI to be semantically well defined in CompCert. For this, we modify the Cminor to remove all semantically ill-defined code. For example, all the local variables which appear in the modules are initialised, we add runtime checks to prevent any division by zero or we verify that function calls match an existing function signature.



```
void main(void) {
  int i = 10;
  int * p = &i;

  *p = 11;
  *&i = 12;
  return;
}
```

```
void main(void) {
  void * sb = malloc(4);
  *(sb + 0) = 10;
  int * p = sb + 0;

  *p = 11;
  *(sb + 0) = 12;
  free(sb);
  return;
}
```

Figure 6: Using `malloc` to allocate memory for the frames (Joshua A. Kroll et al.)

2.3.3 Evaluation of the approach

First of all a strong point of this approach is that it has a small *Trusted Computing Base*. For the CompCert-SFI to be valid it is only necessary to trust that our SFI transformations which were not proven (like the axiomatized *mask* function) does not change the behaviour of the dangerous modules. Moreover we need to trust CompCert, especially the small pieces which are not proven like the proof assistant Coq or the assembler.
Another asset is the absence of verifier in the approach which removes the necessity to develop a different verifier for every architecture targeted. Furthermore the SFI transformations are portable on every architecture supported by CompCert. Cminor is also a high level language, therefore the SFI transformations can benefit from the optimization operations of CompCert which can speed up the transformed module.
The performances obtained during the experiments are similar to a compromise between `gcc -O0` and `gcc -O1` for ARM a and x86-32. If we compare the efficiency with CompCert without SFI we have a decrease of 21.7% on x86-32 and 16.8% on ARM.

However CompCert-SFI also got some downsides. First, the correctness property of CompCert does not apply for concurrent programs. Indeed CompCert does not have an operational semantic for multi-tasking programs which means we cannot ensure the SFI properties in this case. Also the

CompCert-SFI approach forces us to compile the dangerous module to have sandboxed executables. Whereas in the traditional SFI the dangerous modules can be distributed as binaries, then we only need to have the approval of the verifier to execute them without dangers.

2.4 Limits of SFI

We saw that SFI techniques confine jumps in the code differently depending on the addressing mode. For a direct addressing, the generator just rewrites directly the targeted address by replacing the most significant bits with the tag of the sandbox. For indirect addressing we add some runtime sandboxing operations to make sure that the targeted address stored in the register is in the sandbox. There is also another way to make a jump which uses an address stored in the stack: the `ret` instruction. Since `ret` does not use any register we can't use our sandboxing operation here. We could just mask the value stored in the stack but like we mentioned Chapter 2.2.2 this solution is vulnerable to race conditions. Indeed we can't guarantee that the value stored in the stack will not be overwritten by another thread between the runtime check and the execution of `ret`. Therefore the code is weak to Returned Oriented Programing (ROP) attacks which will be presented in detail Chapter 3.1. These attacks overwrite the return addresses located in the stack in order to modify the control flow of the program and possibly execute malicious code.

In the works presented, only NativeClient [9] addresses this issue. The idea is to remove all `ret` instructions and replace them by the composition of a `pop` and a `nacljmp` which is the SFI `jmp`. This simple solution works but the modern processors use numerous optimizations reliant on the use of `ret` instructions. Replacing `ret` instructions greatly decreases the efficiency of the programs. Thus we propose in the following section a more flexible and efficient solution to protect return addresses against the famous ROP attacks and this solution can complete the SFI implementations presented perfectly.

3 Overview of the approach

Many attacks on software aim at diverting the control flow of the targeted program. Among those, *Returned Oriented Programming* (ROP) attacks specifically try to overwrite the return addresses. By doing so the attacked function returns to a malicious piece of code that will get executed. Stack overflow is an example of such ROP attacks. We propose a solution against ROP attacks which combined with SFI would protect from most of control-flow interference attacks. Inspired from SFI techniques we aim to prevent any overwriting of the return addresses. To do so we need to know these return addresses locations in the memory. Therefore our approach consists in modifying the stack structure in order to have a way to distinguish the return addresses locations. With this knowledge we are able to put a mask, as in SFI, before every dangerous write instruction and prevent any ROP attack.

3.1 ROP attacks

We want to protect our program against ROP attacks. These attacks are directed against the stack and especially the return addresses located in the stack. We begin by a short introduction about the mechanisms behind the stack. Then we explain how ROP attacks work with the example of a classical buffer overflow.

3.1.1 The stack

The stack is a specific area of the memory of a program. The memory allocated to a program is divided among multiple areas like the stack (which we are going to detail), the heap (where we put dynamically allocated or global variables) and the code is also allocated in a specific area. The stack is composed of frames and each of them are related to a function being executed. Frames are piled up on the stack following the FIFO rule (*First In First Out*). Explicitly, every time a function is called, a new frame is created and placed on the top of the stack. Reciprocally when a function terminates its frame will be popped out of the stack. Frames contain multiple kind of data related to their function like local variables, parameters of the function and return addresses. Return addresses indicate the point of execution to return to when a function terminates. When popping a frame the program is supposed to execute the code at the address matching the value contained in the return address. ROP attacks aims at overwriting these return addresses which enables them to execute malicious code hidden in another part of the memory instead of continuing the normal flow of the program.

3.1.2 Buffer overflow

Stack overflows are the most popular ROP attacks ref . In Figure 7 we can see an example of buffer overflow written in C. The goal of this code is to execute the function called *evil_code()* which just prints "Argh, we got hacked!\n" line 6 of Figure 7. The function *evil_code()* should not be executed in our program, we suppose it is never called. The code was compiled with *gcc -m32 -fno-stack-protector* to remove all stack protections used by gcc. The output of the code from the successful buffer overflow can be seen Figure 8.

We printed out the stack before and after the attack to have a better comprehension of the attack. We see Figure 8 the consequences of the buffer overflow in red. The stack was overwritten

```
1  void evil_code() {
2          printf("Argh, we got hacked!\n");
3  }
4
5  void foo(char* input){
6          char buf[1];
7          ... code ...
8          strcpy(buf, input);
9          ... code ...
10 }
```

Figure 7: Example of buffer overflow in C

and the return address was modified to the address of *evil_code()* which code was successfully executed.

The vulnerability resides in the function *strcpy* line 8. *strcpy* just copies characters one by one until it finds "0"(which corresponds to the end of a string) in the source string. However our source string can contain many more characters than *buf* is supposed to have. Indeed *buf* is declared line 6 as an array of 1 character and our source string is the argument that we give to the program. If the source string is bigger than the destination then *strcpy* just continues to write the source string over others variables location in the stack and possibly reach the return address. The variable *input* has the value of the parameter we give to our program and is the string we are supposed to copy in *buf*. After few tries and fails we found the correct parameter to successfully do the buffer overflow. This input can be seen on the first line of Figure 8 which is *python -c 'print 13*"a"+"\x7b\x84\x04\x08'* or *aaaaaaaaaaaaa\x7b\x84\x04\x08*
In our example we filled the stack with "a" (which corresponds to "61" in ASCII) until we reached the return address. We can see the consequence of the attack in the output Figure 8, where the stack is full of "61" after executing *strcpy*. When we reached the return address we overwrote it with the address of *evil_code* which was `0x0804847b` given on the second line of Figure 8. This way, the next instruction that is executed after *foo* finishes is the function *evil_code*. At the end of the program we can see that we get a *Segmentation fault (core dumped)*, which is normal because we messed up the stack when we overwrote is with "a". But since we managed to execute *evil_code* the attack is still successful.

Even though the buffer overflow was a basic one, it still happens to see the usage of vulnerable functions like *strcpy* in the industry. Furthermore there exist much more sophisticated ROP attacks which are very effective as witnessed by security vulnerability reports. Before, classic ROP attacks tried to execute some malicious code that was inserted manually in the stack or in the heap. But nowadays most system render their stack and their heap non executable which prevents the classic ROP attacks. However modern ROP attacks are now able to create malicious code from different pieces found in the program or in the libraries used. The most famous one is called *return-to-libc* [2][7] attack which uses pieces of code from the *libc* library to create malicious code. Then the attacker still try to deviate the control flow of the program through return addresses to execute the malicious code just assembled with *libc* code.

```
terminal$ ./buffer $(python -c 'print 13*"a"+"\x7b\x84\x04\x08"')
Address of evil_code = 0x0804847b
Stack before:
0xf7712000
0xff957998
0xf7593d26
0xf7712d60
0x0804868c
0xff957978
0xf7593d00
0xf7713dc0
0xf77828f8
0xff957998
0x08048510              //Return address of foo

Stack after :
0xff958161
0xff957998
0xf7593d26
0xf7712d60
0x0804868c
0xff957978              //Buffer overflow
0x61593d00              //"a"
0x61616161              //"aaaa"
0x61616161              //"aaaa"
0x61616161              //"aaaa"
0x0804847b              //"\x7b\x84\x04\x08", evil_code address

Argh, we got hacked!    //Success! evil_code was executed
Segmentation fault (core dumped)
```

Figure 8: Output from buffer overflow example

3.2 Description of the approach

We want to protect programs against ROP attacks like the buffer overflow seen previously. For that, we need to prevent any return address from being overwritten illegally. The only moment they should be written over is during a function call routine. Our idea is to be able to add runtime checks in the code like SFI, therefore we need to be able to check if an address is the location of a return address.

The biggest difficulty is to be able to know if a location in the stack corresponds to a return address or not. Indeed the stack grows through function calls which pile up stack frames. These frames are constructed dynamically depending on the function, hence the locations of return addresses are unpredictable. As it is we do not have enough information to correctly protect return addresses since we do not know precisely where they are located.

Several solutions exist for this issue. We could, for example, add a lot of meta-data during the

compilation to have extra information and then effectively protect the return addresses. Another solution is to create a second stack called *shadow stack*. We would then have complete control over the *shadow stack* which allows us to defend against ROP attacks.

Our solution is to modify the current stack structure to be able to know the return addresses locations easily. The main idea is to fix a constant offset n between return addresses allowing us to exactly know where a return address is located relatively to the others. We will explain thoroughly the approach that we want to apply in the following section.

3.2.1 Proposed solution

Fixing return addresses locations and stack alignment. We want to be able to decide if a pointer points to a return address at runtime. With this knowledge we will be able to detect if an instruction may compromise our program. The idea is to modify the stack layout in order to have a constant offset n between neighbouring return addresses. This way we know that the neighboring return addresses are always located at a distance n from a frame return address. Furthermore all the other return addresses are separated by a distance which is necessarily a multiple of n. For example suppose that we know the location of a return address, we call this location c. Since all neighboring return addresses are separated by the exact amount n, we know that the following return addresses locations will be $c + k * n$ with $k > 0$. Reciprocally the previous return addresses will be located at $c - k * n$ with $k > 0$.
Thus we can generalize the two previous relations and for every return address location a we will have the relation $a \ mod \ n = c$.

The next hurdle is to choose n and c cleverly. For n, the most important thing is that frames should have enough space to store all the needed data. Therefore we define the value of n as the biggest frame size of all the functions in a program. If the return addresses are separated by this amount we are sure that every function will have enough space in the stack for its frame.
Afterwards we have to define c. The best way that we found is to modify the stack in order to have the first return addresses location to be equal to c. If we are able to do such a thing, we can also easily define the value of c and for simplicity we would like to have $c \ mod \ n = 0$

The Figure 9 pictures the transformation we want to apply to the stack. On the left we have represented an usual stack with return addresses all over the place. Since these locations are unpredictable it is really difficult to pinpoint their location. After transforming the stack (stack on the right) we can see that the different addresses are separated by the same constant n.

We have two relations about return addresses location:

$$[\forall a \in Ret\_locations, \ a \ mod \ n = c] \text{ and } [c \ mod \ n = 0]$$

Ret_locations is the set containing all the locations of return addresses. We can combine our two properties and we obtain:

$$\forall a \in Ret\_locations, \ a \ mod \ n = 0$$

Detection of dangerous instructions. The second step is to detect every instruction possibly harmful to return addresses. We consider as dangerous every instruction that can freely write to the

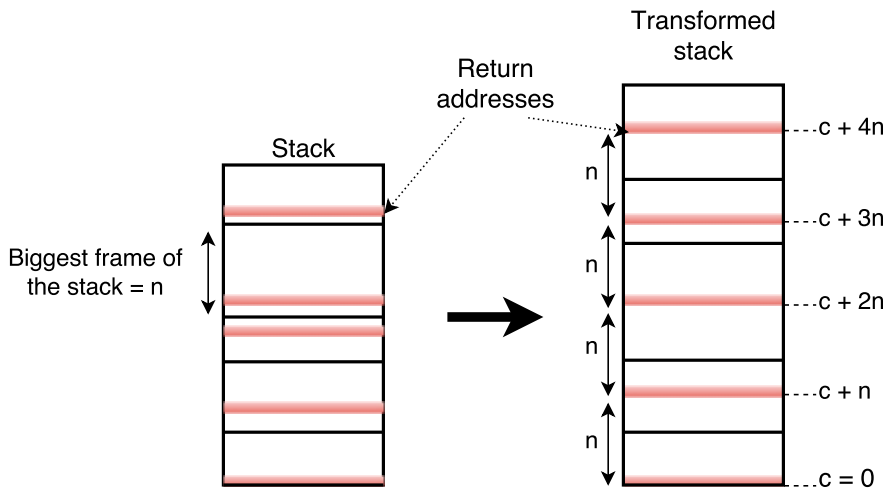


Figure 9: Stack modifications

memory. Our approach is mainly related to the C language. In C, instructions that fit such criteria are assignment to pointer dereference in the form of *(pointer)* = value* or *(pointer+offset)* = value*. In the previous example of buffer overflow Figure 7, the vulnerability resides in the function *strcpy* line 12. To pinpoint the dangerous instruction let's check the source code of *strcpy* Figure 10 from Apple. We can see in the *while* loop line 10 that *strcpy* copies characters one by one from the source string *s2* to the destination *s* until it finds a character equals to *0*. To copy the characters, *s* and *s2* are pointers which initially point to the memory area of the destination and the source string. Then until it finds a character matching *0* the value pointed by *s2* is copied to the location pointed by *s* and the pointers are incremented. The harm happens when the source *s2* is much longer than the destination. In this case we continue to copy to the location pointed by *s* even if the memory written to does not belong to the destination string anymore.
In this example we see clearly that it's the pointer dereferencing that allows one to write directly in the memory. For that reason we target such type of instructions in our approach.

```
1 char * strcpy(char *s1, const char *s2) {
2         char *s = s1;
3         while ((*s++ = *s2++) != 0)
4                 ;
5         return s1;
6 }
```

Figure 10: *strcpy* source code from Apple

Securing dangerous statements. Finally when we have detected all the dangerous statements we transform the module code. Before each of these dangerous statements we add a protection mechanism similar to sandboxing in SFI. The algorithm of the runtime check is represented Figure 11:

1. We check if the address is in the stack. Return addresses only exist in the stack, we do not need to concern ourselves with the other accessible memory area: the heap.
2. If the address is in the stack we check if the target address a verifies our equality $a \ mod \ n = 0$. If it does then it's a return address location.
3. If a target address abides by the two previous condition, it's an illegal instruction and we make the program crash. If it does not then the program just continue to run like normal.

We want our implementation to respect the property of transparency, if a program is safe then our transformation does not modify its behaviour.

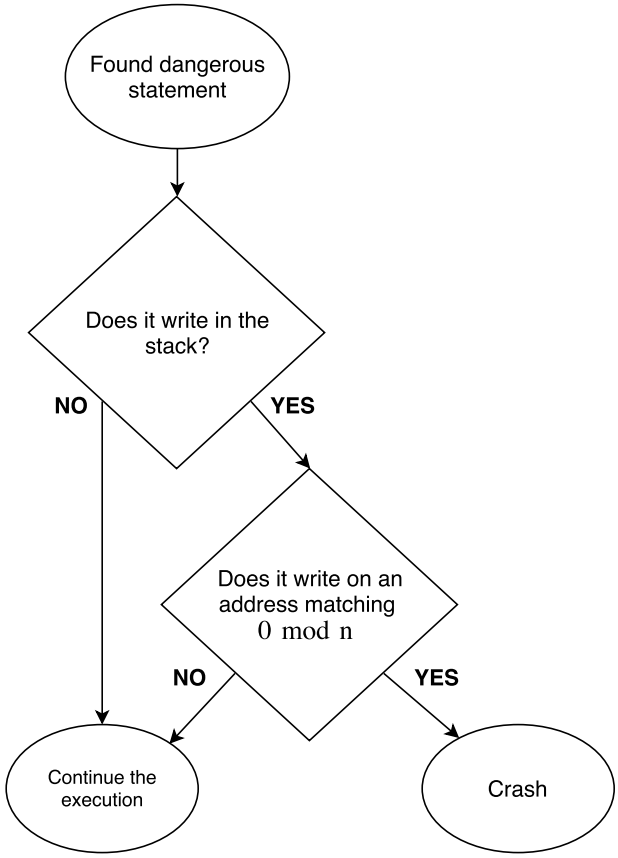


Figure 11: Runtime check algorithm

To sum it up, our approach aims at having an easy way to know return addresses location and then adds a runtime check before every dangerous instruction to prevent illegal writing on return addresses location. To do this we divided the approach into four phases:

1. Fix stack frames size
2. Align the stack

3. Detect dangerous statements

4. Secure the dangerous statements

3.3 Security properties

The approach we propose is composed of four phases, to get the confidence that our idea is effective in protecting return addresses we formalize the properties we expect from each phase. Furthermore like we pointed earlier, we work with the certified compiler CompCert. The ideal way to be sure of our idea would be to prove it with Coq the proof assistant, the language used to build CompCert. By working with these tools we hope that eventually we will be able to prove some security guarantees brought by our approach.

1. **Fixed stack frames size**
    - Return addresses locations are all separated by a constant offset n bigger or equal to any frame of the stack
2. **Stack alignment**
    - The first return address location c of the program verifies $c \ mod \ n = 0$
3. **Detection of memory write statements**
    - Every statement of the analysed code that might modify the stack memory state is detected
4. **Securing memory write statements**
    - The protection triggers an error behaviour if the code try to write on a protected address

Combined, properties 1. and 2. give us the guarantee that all the return addresses locations verify $a \ mod \ n = 0$ with n the fixed offset between return addresses. Basically we make it so the protection mechanism prevents any write on addresses located in the **stack memory area** and verifying $a \ mod \ n = 0$.

Another property we did not mention yet is that all our transformations need to be transparent. In other words, if we apply our methods on a program which is already safe then its behaviour is not affected. We explain how we ensure this property more thoroughly in Chapter 4.

For these properties to always be true we need to place some conditions which we are going to list in the following section. Our approach guarantees that if all the properties mentioned are fulfilled then the program is protected against ROP attacks.

3.4 Analysis of the approach

3.4.1 Conditions

The solution we have just presented can bring very strong security properties against ROP attacks. However for this approach to work we need certain hypothesis to be true. Indeed some of the properties enumerated before are invalidated after certain operations.

Stack modifications. Every operation that disrupts the stack structure may nullify our property that says "every return addresses are separated by a fixed offset". For example x86 architecture use the ESP register to keep track of the stack growth. If we fiddle with it we may introduce a shift in the return addresses locations. Then our runtime check of the return addresses locations property $a \ mod \ n = 0$ would not be relevant anymore. For example, Figure 12 shows a piece of inline assembly which disrupts the stack line 2. Inline assembly allows one to put some assembly code in the middle of C code. Here the assembly decrements the stack pointer ESP. By doing so the stack is shifted by an amount of 50 bytes and our formula for the locations of return addresses is not correct anymore.

Insecure libraries. For our approach to work we need to have all dangerous write statements to contain our runtime checks. Hence all executed code must have been compiled with our transformation. For example, the C library *glibc* contains multiple insecure functions like *printf, strcpy...* Furthermore those flawed functions are common vulnerabilities for *buffer overflows* attacks which are a type of ROP attack. To avoid this issue we need to rewrite the glibc or compile it with our tools.

Modules need the same offset. If a program uses multiple modules or libraries, they need to be compiled with the same offset n. Indeed if the offset of the different modules are different we cannot use the previously defined relation $a \ mod \ n = 0$ anymore. Thus it is not possible to easily know if a location corresponds to a return address.

```
1 int foo(int a) {
2         asm(``\$sub 50, \%esp'');
3         //This line does the operation ESP = ESP - 50
4         //This disrupts the stack layout we establish in our
              transformation
5         printf("Hello world!");
6 }
```

Figure 12: C inline assembly

3.4.2 Discussion

We have presented the principle of our approach in this chapter. Then we mentioned some necessary conditions for our solution to work properly. In this section we are going to discuss about the pros, cons or remarks about the proposed solution.

The benefits of our transformation is clear, any code compiled with a compiler enforcing our methods is unable to interfere with the control flow of our program through return addresses. Furthermore if we combine our solution with the SFI presented earlier we can have some strong security properties on the execution of dangerous modules with our main program. In return our approach may also have some impact on portability and efficiency:

Architecture dependant. Our solution depends a lot on the stack layout of the program. Indeed fixing the size of the frames requires us to modify the original stack layout. Therefore since the stack layout varies depending on the architecture and compiler you are using, the modifications that have to be done are also different. We can then easily understand that we need a different implementation for every existing stack layout. Moreover since these layouts can be really different it might be quite complicated to implement our solution on certain of them. In the implementation we present after we focus solely on x86-32 architecture with the compiler CompCert.

Memory consumption. Since we are fixing the size of the frames instead of having dynamic sizes the memory usage of the stack is bigger. We have the issue of choosing an adequate size for the frames in our solution. The easiest one is to take the maximum frame size of the program as the constant size for all the frames. The downside is that we might have a memory usage explosion from our stack. We did not encounter any issue about memory during the tests we did but the impacts may be visible on especially big programs. It might be interesting to study the efficiency of our approach on the growth of the stack.

Despite the cons presented we believe the benefits we gain from this method is worth it. We are going to present in the following section the implementation we made based on the ideas we introduced here. This implementation was made with the compiler CompCert for the x86-32 architecture. We are targeting programs written in C, which explains that all the examples we used were related to the C language.

4 Implementation

For the implementation we work with the compiler CompCert. CompCert already has an implementation of SFI presented earlier. Thus if we could combine our approach with the SFI, any program compiled with CompCert would have strong security guarantees. Furthermore CompCert is written with the Coq proof assistant, our implementation then paves the way for a formal proof of security properties. In this section we explain in detail how to enforce our approach and the different choices we made during the process.

Our approach is separated in four phases: "Fixed stack frames size", "Stack alignment", "Detection of memory write statements" and "Evaluation of the implementation". We are going to detail the implementation of these phases in the following sections. These transformations are deeply linked to the stack layout, hence to have a better understanding we are going to start by introducing the CompCert stack structure.

4.1 CompCert stack

The layout of the stack is dependent of the architecture and the compiler/interpreter used. For the sake of comprehension of the future sections we describe here the stack layout of x86-32 in CompCert. The stack layout of CompCert x86-32 is pictured in Figure 13. First of all we can notice that the stack grows downwards, it means that the stack grows from the highest addresses to the low ones. As we can see the usual data are stored in this stack like local variables, parameters, register states and the return address.

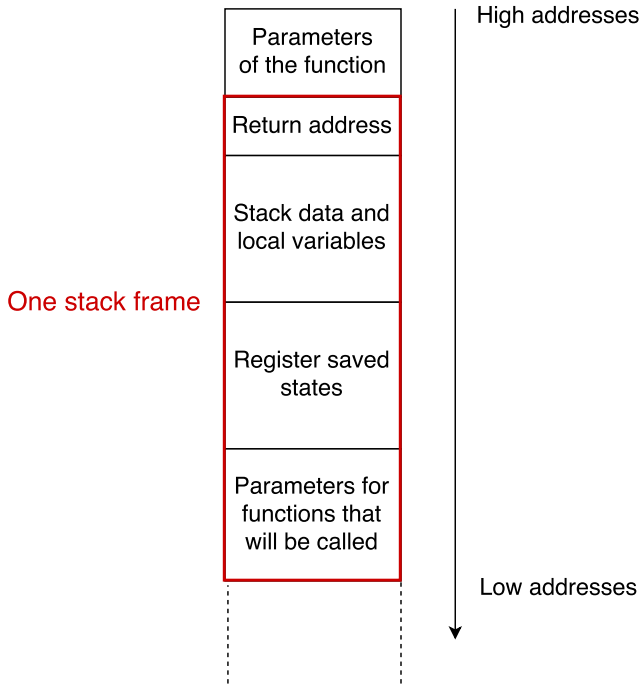


Figure 13: CompCert x86-32 stack layout

Each frame is built when a function is called, the different steps related to the creation of a

frame is called *function call routine*. CompCert function call routine is described in the Figure 14. Each phase of the function call routine of the Figure 14 is explained just here:

1. Write the return address
2. Allocate enough memory for the rest of the stack frame
3. Save registers states in the stack
4. Execute the function body (use the memory for local and stack data)
5. When calling another function, place its parameters at the end of the stack and repeat the process

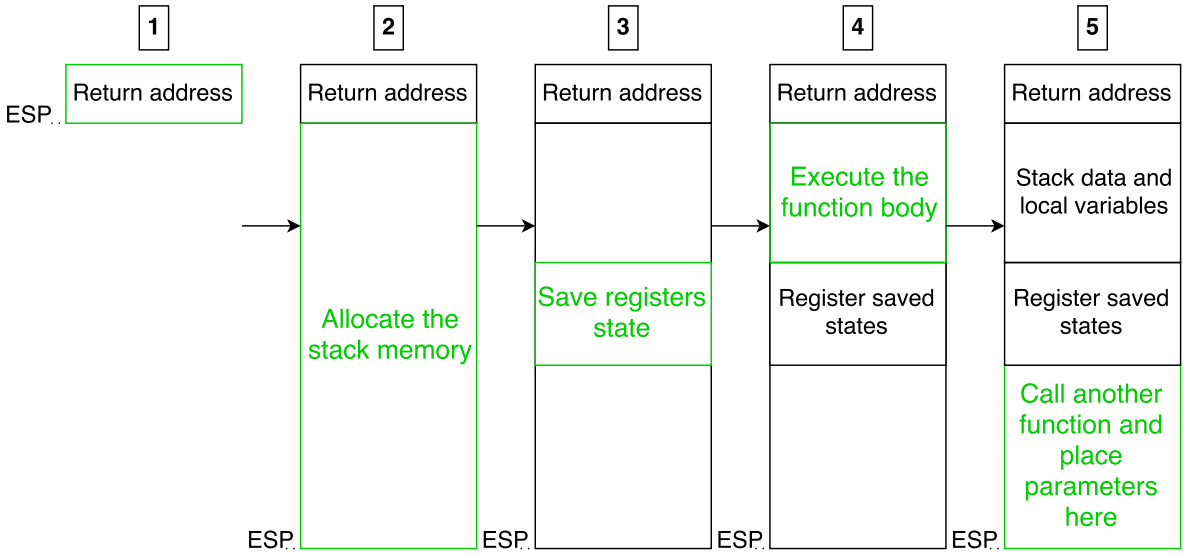


Figure 14: CompCert function call routine

When returning from a function, the return routine is pretty much the opposite:

1. Restore registers state
2. Deallocate the stack until the return addresses
3. Pop the return address memory and jump to the value stored in it

4.2 Fixed stack frames size

During this phase we want to ensure these two properties:

- Return addresses locations are all separated by a constant offset n bigger or equal to any frames of the stack

Fix the frames size. Fortunately in the function call routine of CompCert the return addresses are always at the top of their frames. This particularity makes the task easier, indeed, since the location of the return address is fixed in the stack we can simply fix the size of the frames to have a constant offset between the return addresses.
To modify the stack layout of CompCert we just need to modify the record describing the frames presented Figure 15. In our case the record has different fields like the size of the whole frame and the offset of the different data of the stack. Hence we give to the field it *fe_size* line 2 a constant $fe\_size = n$ which will be a power of two bigger than its previous value. We also need to keep the return address at the bottom of the frame (the stack grows downwards). Thus we have $fe\_ofs\_retaddr = n - 4$ line 4 because the return address is four bytes long. The locals and stack data will then naturally fill the extra space left by the shifted return address like shown Figure 16.

```
1  Record frame_env : Type := mk_frame_env {
2    fe_size: Z;
3    fe_ofs_link: Z;
4    fe_ofs_retaddr: Z;
5    fe_ofs_local: Z;
6    fe_ofs_int_callee_save: Z;
7    fe_num_int_callee_save: Z;
8    fe_ofs_float_callee_save: Z;
9    fe_num_float_callee_save: Z;
10   fe_stack_data: Z
11 }.
```

Figure 15: CompCert x86-32 frame record

In Figure 16, we present the transformation of the CompCert frame to a new frame with a fixed size equal to n. We can see that the stack data and locals part have increased and filled the empty space further in the bottom.

Choice of frames size. We chose to limit the choice of the frames size to powers of two. Indeed, since addresses are written in binary having a power of two as constant offset will ease the runtime checks implementation. By using powers of two, all the return addresses will have the same least significant bits. For example if the first return address location is `0xffffff910` (in hexadecimal) and our constant frame size is 2^8 = `0100` (in hexadecimal). Then all the following return addresses locations will be `0xfffff810` (the stack grow downwards), `0xfffff710`, `0xfffff610`... And their least significant bits are actually always equal to `10`. This particularity will help during the phase of runtime checks that we will explain in Section 4.5.

To prevent having compiled programs with too small stack frame, we added in CompCert a check. This test verifies that the chosen size is bigger than any dynamically calculated one. If not, the compilation fails. This way the chosen size corresponds to the smallest power of two which is bigger or equal to any dynamically calculated frames size of the program.

We can see in Figure 17 the effect of our implementation. The left stack is the usual layout of CompCert stacks with the return addresses located at the top of the frames. We call F_{size} the size of the biggest frame of the whole program.

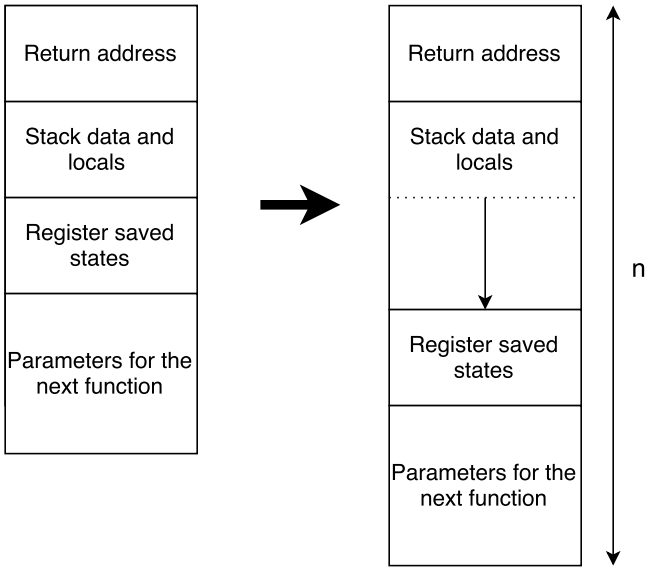


Figure 16: Transformation of CompCert frame

For our transformed stack we have to chose a fixed size for the frames and it needs to be a power of two, bigger or equal to F_{size}. In Figure 17 we chose 2^8 to continue the examples we gave before. We can see that the stack on the right has fixed size frames equal to 2^8 and the return addresses are all separated by the same offset dues to CompCert stack layout. The implementation effectively fulfills the property of having constant offset between return addresses. Furthermore we can see that the location of the return addresses are `0xfffff910`, `0xfffff810`, `0xfffff710`... Hence all the return addresses have the particularity of always having the same least significant bits (`10`). This particularity will be used later for the implementation of the protection mechanism.

4.3 Stack alignment

The implementation in this section has to do a transformation that makes the following property true:

- The first return address location c of the program verifies $c \ mod \ n = 0$

ASLR. This property was hard to make true because of the existence of ASLR that we talked about before. Indeed since there is randomization of the stack addresses we needed to find an operation that could align the stack correctly for any initial location of the stack. ASLR randomizes partially the addresses, for example the beginning address of the stack always has the following format `0xff*****0`. ASLR keeps the eight most significant bits so the different memory areas like the stack are still contained in their reserved area. The four least significant bits of a new frame are always `0` to keep a certain alignment in the memory which increases the efficiency.

Stack alignment. We had multiple choices for the implementation of this property. One of them was to modify the main function of the protected program in order to align the return addresses

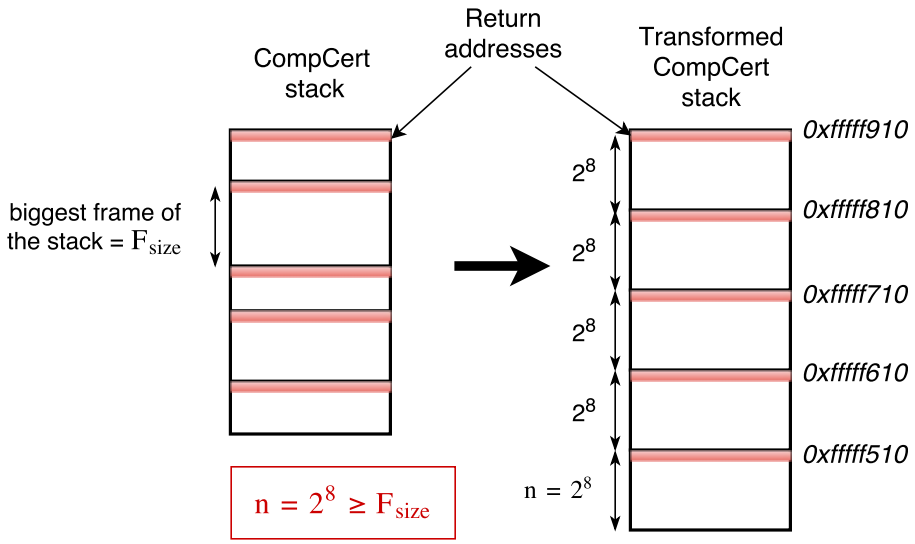


Figure 17: Fixing the size of CompCert stack frames

locations correctly. Another idea was to modify the prelude of a program, the prelude is a piece of code created by the compilers which is executed before the program. It is necessary for any program to have this prelude to work correctly.
Eventually we chose to introduce before the *main* function of the program an artificial *main*. Its role is to align the stack in order to make all the incoming return addresses locations to match our formula $a \ mod \ n = 0$.
Since we have to modify the stack structure we did our transformation at the assembly level(ASM) of CompCert. Indeed the stack pointer ESP which is responsible for the stack growth is only available in ASM. ASM is the lowest level before binary code, though it is difficult to modify ASM correctly since you have to manipulate low level objects. By creating a separate artificial *main* function we avoid taking the risk of bugging the prelude or the program's *main* function.
This approach has one definite advantage over the others solutions listed before. By introducing an artificial *main* we make a clear separation with the code of the program like the original *main* and the prelude. Combining the fact that the original *main* and the prelude might have complicated behaviour and that we do the transformation at a low level. It is much safer to create a whole new function which can help us avoid introducing bugs in the assembly. Hence the whole artificial *main* function needs to be written by hand in ASM. The code includes the function routine and the call to the original *main* of the program.

Figure 18 represents the stack alignment transformation. The left stack is CompCert stack with fixed frames size equal to $n = 2^8$ like we had in the previous section. From this stack we show the consequences of our operation. We insert before the *main* function of the program an artificial *main*. Thus the frame of this artificial *main* is the first frame of the whole stack. The artificial *main* objective is to align the stack in order to have the next return address location a equal to $a \ mod \ n = 0$.
We can see on the left stack that the return address of *main* was previously at the address `0xfffff910`. After the stack alignment transformation (right stack) it is located at the address

`0xfffff700`. Since the frames size remain constant we now have all the following return addresses locations verifying $a \ mod \ n = 0$. This was the objective of the whole stack transformation which is now completed. The downside of this implementation can be seen clearly on the Figure 18. Indeed all the return addresses locations verify $a \ mod \ n = 0$ except the return address of the artificial *main* we introduced. Since our approach aims at protecting the locations verifying $a \ mod \ n = 0$, this return address is vulnerable. Nevertheless, to reach this location an attacker would need to either know the exact location either overwrite the whole stack. Both ways are difficult to implement due to ASLR and the fact that we protected the stack from being overwritten at the return address location. We can also come up with a more robust implementation of the artificial *main*. Since we build the ASM code by hand we are not forced to follow the usual function routine and we can have a different and robust frame structure for the artificial *main*.

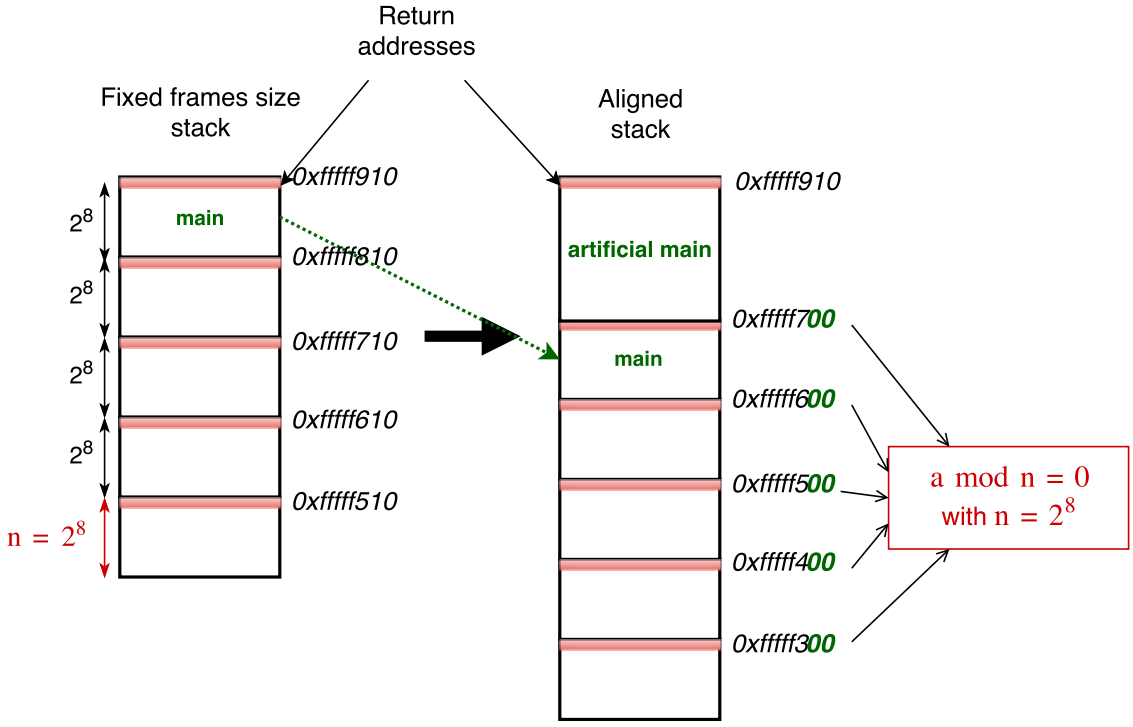


Figure 18: Aligning CompCert stack frames

Alignment algorithm. We present in Figure 19, the algorithm used to calculate the right size in order to have the next return addresses aligned even with ASLR. The algorithm is written in pseudo assembly code.

The easiest way to understand it is to go through it with an example. On the previous examples Figure 17 and Figure 18 our stack starts at the value `0xfffff910`. Hence for the continuity of the examples we keep this value. In Figure 18 we want to have our next frame starting at the address `0xfffff700`. We go through the different lines of Figure 19 to explain the algorithm. *next_frame* corresponds to the value of the stack pointer which gives the starting address of the following frame. The aim of the alignment algorithm is to verify the property $next\_frame \ mod \ n = 0$ for any initial

value of *next_frame* given by ASLR. For the example we will take *next_frame* = `0xfffff840`.

1. we copy the initial *next_frame* location into a register called *reg*.
   The operation is then *reg* = *next_frame* = `0xfffff840`
2. *reg* = *reg* & $(n-1)$ = `0xfffff840` & `0x000000ff` = `0x00000040`
   In our examples we have $n - 1 = 2^8 - 1$ = `0x000000ff`
3. *reg* = *reg* +$(n-4)$ = `0x0000040` + `0x000000fc` = `0x00000140`
4. *next_frame* = *next_frame* − *reg* = `0xfffff840` − `0x00000140` = `0xfffff700`
5. We start the function call routine, here we save registers state in the stack
6. We store the parameters of the original *main* in the stack
7. We call the original *main* function, its frame will start at the location stored in *next_frame* = `0xfffff700`

A small remark on the example is that our algorithm only works if the last bit of the first value of *next_frame* = `0xfffff840` is `0`. However this particularity is not an issue because all compilers follow this rule since it improves the speed of execution due to alignment property.

```
1 move    reg          next_frame
2 and     reg          n-1
3 add     reg          n-4
4 sub     next_frame   reg
5 store   regs_state
6 store   parameters
7 call    main
```

Figure 19: Alignment algorithm

4.4 Detection of memory write statements

Clight implementation. We chose to implement the detection of dangerous statements and also the runtime checks of those statements at the Clight level. This choice is explained by the fact that Clight is a high-level language in the compilation steps of CompCert (it is the closest to C so the syntax is really similar). Indeed, doing our transformations at a high-level is much easier since all the complex compilation operations are done later in the process. For example by using Clight we do not need to bother with low-level objects like registers which if misused can modify the program unexpectedly. Furthermore Clight is a compilation step placed before any optimization of CompCert. This means that our implementation can be optimized automatically by CompCert thus improving our performance.

Clight semantic. We have to make sure that we cover all possibly harmful statements with our runtime protection. Since we are working with the compiler CompCert we are going to take advantage of it. CompCert has multiple compilation steps which have all been proven from C to assembly language. To make these proofs, a semantic was defined for each language of the compilation. The semantics all relate to the memory model briefly described in Section 2.3. To detect all dangerous statements we looked at the semantic of Clight and found all statements that in the memory model could write freely in the memory.

```
1  Inductive statement : Type :=
2    | Sskip : statement
3      (**r do nothing *)
4    | Sassign : expr -> expr -> statement
5      (**r assignment [lvalue = rvalue] *)
6    | Sset : ident -> expr -> statement
7      (**r assignment [tempvar = rvalue] *)
8    | Scall: option ident -> expr -> list expr -> statement
9      (**r function call *)
10   | Sbuiltin: option ident -> external_function -> typelist -> list
       expr -> statement
11     (**r builtin invocation *)
12   | Ssequence : statement -> statement -> statement
13     (**r sequence *)
14   | Sifthenelse : expr  -> statement -> statement -> statement
15     (**r conditional *)
16   | Sloop: statement -> statement -> statement
17     (**r infinite loop *)
18   | Sbreak : statement
19     (**r [break] statement *)
20   | Scontinue : statement
21     (**r [continue] statement *)
22   | Sreturn : option expr -> statement
23     (**r [return] statement *)
24   | Sswitch : expr -> labeled_statements -> statement
25     (**r [switch] statement *)
26   | Slabel : label -> statement -> statement
27   | Sgoto : label -> statement
```

Figure 20: Clight statements

In Figure 20, we have exposed all the Clight statements. Among them we are going to focus on the ones that change the state of the memory. When looking at the semantic given to these statements, only four of them can change the state of the memory: *Sassign*, *Sbuiltin*, *Sreturn*.

- ***Sassign***, are used to assign value to variables, it could be considered as an equivalent of "=" in C. These statements will be targeted by our approach.

- ***Sbuiltin***, are used to call builtin functions, which are functions created by CompCert that will be expanded later in the compilation. These statements call functions we trust, that is why we will not consider them as dangerous. We could also look at the builtin functions and modify their code to make them safe.
- ***Sreturn***, these statements invoke the function call routine. They are also trusted statements, we will not need to add runtime checks on them.
- ***Sset***, could have been problematic. However they are only used to set value to temporary variables which means that they can't freely access the memory.
- The other statements are all kind of control flow statements, they do not change the memory state.

Among all the statements, our security checks will only apply to the *Sassign* statements. Furthermore we can limit ourselves to *Sassign* statements which left expression can write directly in the memory. Left expressions are the expressions explicitly on the left of the "=" symbol, they are the only expressions which enable us to write in the memory ("Sassign left_expr right_expr" $\leftrightarrow$ "left_expr = right_expr"). The left expressions targeted are then mostly pointers dereference. To be sure that we have all the dangerous instructions, we reiterate the same approach and we take a look at the semantic of the left expressions in Clight.
After checking the semantic of the left expressions, only two types of left expressions are able to reference a location in the memory.

- ***Ederef***, as we predicted these expressions dereference pointers and will be targeted by our approach.
- ***Efield***, they refer to fields of structure and can also point to locations in the memory. These expressions will also be secured with runtime checks.
- ***Evar*** are the last left expressions possible. *Evar* are tied down to a location in the memory. They can't reach return addresses locations hence they are not threats.
- The other expressions are not left expressions which means they do not write in the memory and they are not dangerous in our case. Unary operators *Eunop*, binary operators *Ebinop*, constants *Econst_int Econst_float* are examples of such expressions.

We finally have defined the profile of the dangerous statements that have to be targeted by our approach. To sum it up, the targeted statements are all the *Sassign* which left expression is either *Ederef* or *Efield*.

Now that we can detect the dangerous statements we will now add the runtime checks in the Clight code which will terminate our implementation.

4.5 Securing memory write statements

We want to add a protection mechanism which prevents any dangerous statement from writing on a return address location. These return addresses locations have two special traits:

1. they only exist in the memory area of the stack

2. thanks to our previous modifications, their locations a verify $a \ mod \ n = 0$

If a statement try to write on an address with these two properties then it is an illegal execution and our mechanism will trigger an error behaviour like crashing. The Figure 11 explained earlier can be used as a reminder of the principle of these runtime checks.

Distinguish stack and heap addresses. A program can only use the stack and the heap to store data during runtime. Since return addresses are only present in the stack we need a way to differentiate stack and heap addresses. In x86 architecture, the stack usually grows downwards in direction of the heap. Therefore the addresses of the stack occupy the high addresses and the heap the low ones. The idea is to divide the memory space for the program data into two distinct part. The high addresses for the stack and the low for the heap.
We defined the clear separation at the address `0xff000000`. Every address bigger than `0xff000000` is considered as part of the stack. Reciprocally every address smaller is part of the heap. However this assumption is not always true, in extreme cases the stack or the heap might overflow over this limit. To ensure that our criteria becomes a truth we propose a solution in the future works section Chapter 5.3.

Figure 21 presents an implementation of our runtime checks. The condition line 1 is the branch which separated addresses from the stack or the heap. Heap addresses can't be return addresses location therefore they do not continue the verification process.

```
1 if (targeted_address > 0xff000000) {
2         temp_var = targeted_address & (n-1);
3         if (temp_var < 3) {
4                 Error behaviour
5         }
6 }
7 *targeted_address = value;
8 Continue execution...
```

Figure 21: Runtime checks

Check the property $a \ mod \ n = 0$**.** The second step of the runtime check is to see if the targeted address location a verifies $a \ mod \ n = 0$. It corresponds to the code from line 2 to to line 4 of Figure 21.

- At line 2 we reduce the targeted address to its $log_2(n)$ least significant bits (8 if $n = 2^8$). Since n is a power of two just comparing the least significant bits to 0 is equivalent to checking if the address verifies $a \ mod \ n = 0$. For example we know that the return addresses locations of our previous example Figure 18 were `0xffffff700`, `0xffffff600`, `0xffffff500`... We can notice that we only need to check if the last eight bits (the last two digit in hexadecimal `00` in our example) of an address is equal to 0. If it is we know for sure that the targeted location is a return address.
  For a random authorized location `0xffffff7a2` and a return address location `0xffffff700` with $n = 2^8$, we would have to compare `0x000000a2`(authorized) and `0x00000000`(unauthorized) with 0.

- Line 3 we actually see that we do not compare the previous calculated value with 0 but 3. The reason is that return addresses are four bytes long. It means that the space taken by a return address located at `0xfffff700` would be [`0xfffff700 - 0xffff703`]. It is logical for us to protect the whole return address space, that is why we check if the targeted address least significant bits are smaller than 3.

- Line 4 is executed when there is an illegal behaviour. Currently our error behaviour is to make the program crash by trying to write over some protected memory located at the address 0 which triggers a *Segmentation fault*.

- Finally line 7 if we successfully pass the verification we are allowed to write on the dereferenced pointer *temp_var*.

Currently our protection mechanism use two *if ... then ... else* operation in a row to protect the return addresses. The first to check if the address is in the stack and the second to check the equality $a \ mod \ n = 0$. Considering that we might inject our protection mechanism a considerable number of times for a program using a lot of pointer operations. We wanted to try another possible implementation that may improve the performance our runtime checks. Hence we present in the next section an alternative to the classic *if then else* called branchless statement.

Branchless check. We wanted to limit the overhead introduced by our approach by trying another way to make the protection mechanism.

Branchless code allows one to create code with the same behaviour as a classic *if then else* but without creating any branch for the processor. In other words the processor will not need to execute different code depending on the condition, the code will be linear.
The best way to understand it is to have an example and we show the branchless version of implementation of the protection mechanism in Figure 22.

The branchless code is the code located from line 2 to line 6. The main idea of the branchless code is to create a mask from line 2 to line 5 which will be stored in *temp_var*. If the initial value of *targeted_address* is a return address location the mask should be full of 0 so that the masking operation line 6 gives us a null pointer. In the other case the mask should be full of 1 so that the masking operation line 6 keeps the value of *targeted_address*. The first condition line 1 is the same as before which checks if *targeted_address* is in the stack or the heap. Branchless uses a lots of bit arithmetic and low level properties of two's complement representation of negative numbers. We are going to go through every line of branchless code with two different initial value of *targeted_address*: one is a return address location which is illegal and the other is an authorized address.

For both case we have $n = 2^8$ for the chosen frames size. We start with an authorized address and we take the *targeted_address* = `0xfffff7a2`. Since we have a legal address our branchless operations should leave *targeted_address* unchanged.

2. We want to keep the least significant bits with a *and* operator (`0101` & `1100` = `0100`)
   temp_var = `0xfffff7a2` & `0x000000ff` = `0x000000a2`
3. *temp_var* = `0x000000a2` - `0x00000003` = `0x0000009f`
4. A right shift bit operation (`0100` → `0010`) thirty-one times. If the value is negative then the new bits introduced on the left are equal to 1 else they are equal to 0.
   temp_var = `0x0000009f` >> 31 = `0x00000000`

5. bits inversion (`0100` → `1011`)
   temp_var = ∼`0x00000000` = `0xffffffff`
6. *targeted_address* = `0xffffffff` & `0xfffff7a2` = `0xfffff7a2`

The branchless operation is successful, the initial address was an authorized location, therefore *targeted_address* kept its value during the branchless and the assignment line 8 is successful.

```
1 if (targeted_address > 0xff000000) {
2         temp_var = targeted_address & (n-1);
3         temp_var = temp_var - 3;
4         temp_var = temp_var >> 31;
5         temp_var = ~temp_var;
6         targeted_address = temp_var & targeted_address;
7 }
8 *targeted_address = value;
9 Continue execution...
```

Figure 22: Branchless version of the second check

We go through the case where *targeted_address* is a return address locations, we take *targeted_address* = `0xfffff700`. The destination is illegal, the branchless operations should leave in *targeted_address* an illegal value to make the program crash.

2. *temp_var* = `0xfffff700` & `0x000000ff` = `0x00000000`
3. *temp_var* = `0x00000000` - `0x00000003` = `0xfffffffd` (signed representation)
4. *temp_var* = `0xfffffffd` >> 31 = `0xffffffff`
5. *temp_var* = ∼`0xffffffff` = `0x00000000`
6. *targeted_address* = `0x00000000` & `0xfffff7a2` = `0x00000000`

The value of *targeted_address* at the end is zero. Hence the assignment line 8 will make the program crash since zero is a protected memory area. The initial address was a return address location which means that the illegal operation crashed as we wanted.

This way we can notice that we got the same result as with the *if then else* check but without creating any branch. In our case it is not possible to replace the first *if then else* condition because it would even force heap addresses to go through the second check which we do not want.

5 Evaluation of the implementation

We will show first in this section that our implementation is effective. Afterwards we will present different benchmarks we did to measure the cost of our transformations on different criteria like speed or size of the programs. To conclude we will discuss multiple ideas to improve the current approach.

5.1 Evaluation of security

For the example we are going to confront the buffer overflow presented Figure 7 with our implementation. We modified the code of the buffer overflow a little to simplify the test. To make things easier we just added the function *strcpy* in the program's code instead of calling it from another library. This way the vulnerable code of *strcpy* will be compiled with our transformation and be secured instead of compiling the whole external library.

To check the effectiveness of our approach we compile the buffer overflow in two different ways. The first time we compile it with our implementation but we do not activate the protection mechanism. In other words this executable has fixed frames size but we did not insert the runtime checks so it is still vulnerable to the buffer overflow. The second executables is compiled with our whole mechanism and is protected with runtime checks. However the two executables were compiled with the same stack frames size, $n = 2^6$. The output of the two binaries are presented Figure 23 and 24.

Fixed stack frames without protection (Figure 23). Like in Figure 8 we can see that we have printed *Argh, we got hacked!* which means we managed to execute the function *evil_code*. The buffer overflow is successful, we can see that we needed to change the input for the attack to work. Indeed since we changed the compiler and the stack layout the input had to be different to manage the attack.

Complete implementation with protection (Figure 24). For the sake of the example, we enriched the error behaviour statement which now calls "*printf(Address: %p, tmp)*" before crashing. In *tmp* we stored the value of the address for which the protection mechanism triggered the error behaviour. First of all we can see that we made the program crashed before *evil_code* was executed. Our implementation successfully does what we aimed for, we made the program crash on a dangerous program. Furthermore we can even see the address on which the program tried to write and was not allowed to. In this example our protection mechanism crashed when the program tried to write on the address $a = \texttt{0xffa2b0c0}$. Since the frames size is $n = 2^6$ all our return addresses location should have their least six significant bits equal to 0. We check that $a = \texttt{0xffa2b0c0}$ is effectively a return address. First of all $a > \texttt{0xff000000}$ which means that a is actually part of the stack. Moreover we need to verify the equality $a \ mod \ n = 0 \leftrightarrow (\texttt{0xffa2b0c0}) \ mod \ (\texttt{0x00000040}) = 0$ which is true. Both the conditions are fulfilled which means that a is a return address location and that our mechanism properly worked.

5.2 Evaluation of performance

As mentioned earlier, we also want to measure the impacts that our transformation may bring in terms of memory and efficiency. We made some experimentations in order to figure out these

```
terminal$ ./buff_comp_weak $(python -c 'print "a"*44+"\xb0\x84\x04\x08"')
Address of evil_code = 0x080484b0

Stack before:
0xf756e724
0xff9dab44
0xff9dac74
0x00000002
0xf7731000
...
0xf75b2d00

Stack after :
0xff9db132
0xff9dab44
0xff9dac74
0x61616161
0x61616161
...
0x61616161

Argh, we got hacked!
Segmentation fault (core dumped)
```

Figure 23: Output from vulnerable CompCert executable

questions. Furthermore we also had to verify which implementation was the most efficient between the branchless and the classic *if then else*. For our different experimentations we had 74 different C programs which were present within the CompCert project for testing purposes. The performances of our approach are compared with the original CompCert project which performances are similar to `gcc -O1` [4]. The machine used for the benchmarks possesses four 64 bits processors Intel Xeon E5 with a capacity of 2.80GHz and 8GB of memory. The operating system of the machine is Fedora 23. For all our executables compiled with fixed frames size we chose the smallest size possible for the programs to be able to run correctly. Our script just try to compile every program with an incrementing frames size until it manages to complete the compilation process.

5.2.1 Efficiency

We wanted to measure the efficiency of our different transformations and compare them to the original CompCert. During our test we measured four different implementations: the original CompCert, our transformation just with fixed frames size, with fixed frame size and classic *if then else* runtime checks and the branchless version. To calculate the execution time we used the Unix utility *time*. The results are presented in the following graph Figure 25. For more relevant results we removed from all the tested programs the ones which execution time was too short because the Unix utility *time* imprecision could impact the result greatly. The histogram Figure25 shows the percentage of overhead measured by three different implementations compared to the original

```
terminal$ ./buff_comp_weak $(python -c 'print "a"*44+"\xc0\x85\x04\x08"')
Address of evil_code = 0x080485c0

Stack before:
0xf7554724
0xff9a88c4
0xff9a8a04
0x00000002
0xf7717000
0000000000
0xf7598d26
0xf7717d60
0x08048654
0xff9a88cc
0xf7598d00

Address: 0xffa2b0c0
Segmentation fault (core dumped)
```

Figure 24: Output from protected CompCert executable

CompCert compiler for multiple programs. The three implementations are "Fixed frames size" which is our implementation with modified stack layout but without adding any protection. There is also "Secure" which is the implementation with classic *if then else* runtime checks. The last one "Branchless" is the version with branchless checks.

Fixed frames size performance. We notice first that 17 out of the 23 programs an efficiency comparable to the original CompCert (100) for "Fixed frames size". This result was expected, our transformation concerning the stack layout does not involve any extra code it just changes the structure of the memory. However we still have 6 unexpected results, some are faster and some are slower. *spectral* implementation even performs three times slower than the ones with extra code for protection, which seems strange. We suspect for this mysterious cases that it is due to some special alignment in the memory or location to the processor that influence the performance. More investigations would be needed to determine the exact reason of this phenomenon.

***if then else* and branchless comparison.** As for the comparison between the *if then else* and the branchless implementations we were disappointed by the results displayed by the branchless version. Indeed the branchless is clearly faster than the classic *if then else* for only three programs in Figure 25. It is distinctly slower for 5 programs and for the rest the performance are similar. However to really compare the two implementations of our runtime checks we should only compare programs that contains a fair amount of them in their code, otherwise we would not really compare branchless and branching.

Figure 26 presents the number of time our runtime checks are executed. We only selected the programs which results had a significant amount of runtime checks. This graph also separates two cases, tested addresses which were in the heap and addresses from the stack. Indeed in our

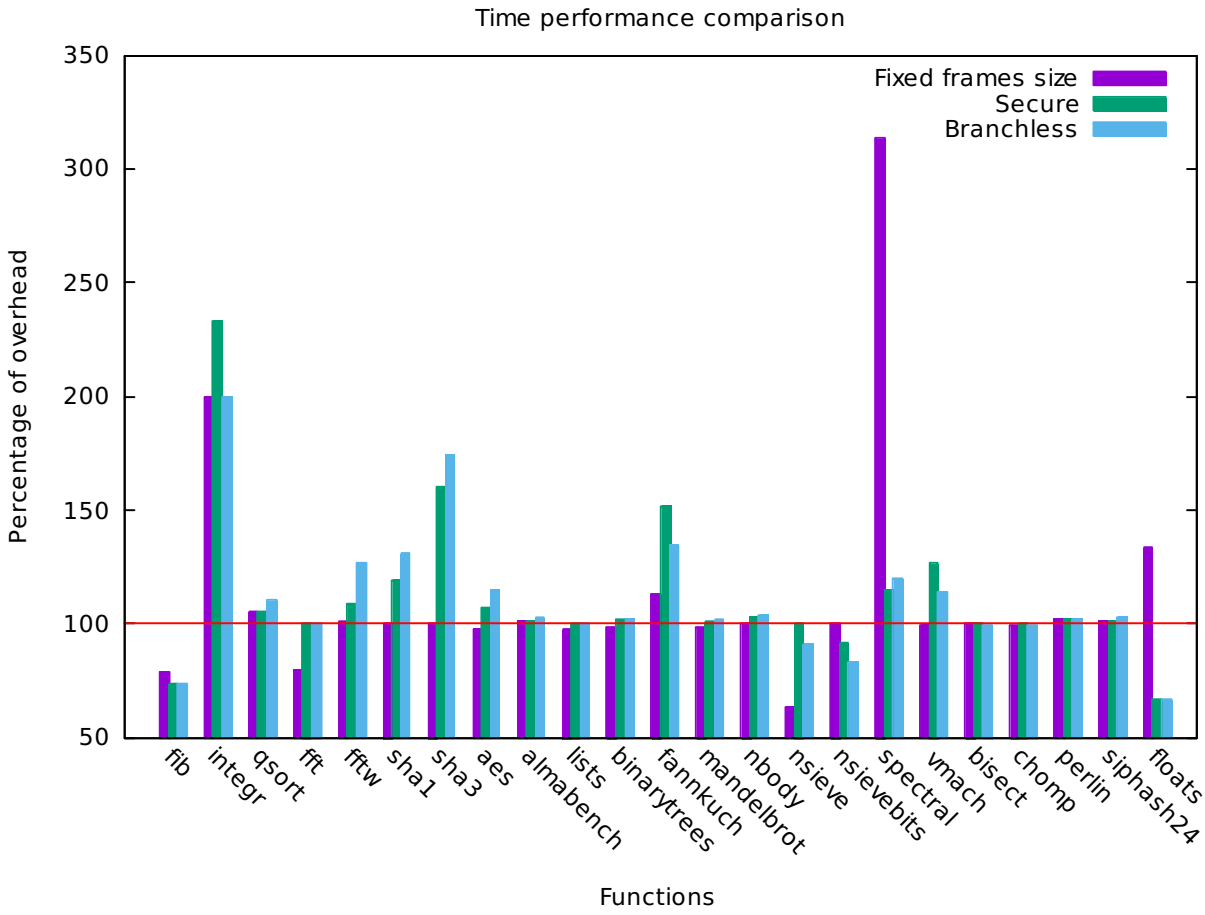


Figure 25: Execution time percentages compared to the original CompCert

implementations addresses from the heap only need to go through one conditional jump whereas for the stack we also check the equality $a \ mod \ n = 0$. Most of the programs which have not been represented had no runtime checks at all because they did not contain any dangerous instructions. The number of occurrences are on a logarithm axis since the values had big gaps between them.

To compare the branchless and *if then else* we will limit ourselves to the programs which executed more than $1 * 10^8$ runtime checks which gives us the programs *sha1*, *sha3* and *aes*. In Figure 25 these three programs branchless version are distinctly slower than the *if then else* version (only 5 programs out of 23 show this particularity). Even though the number of programs analysed is low there is a high possibility that for our implementation the branchless runtime checks are actually slower than the *if then else* one.

What can makes branching instructions expensive in term of time is due to instructions pipelining. The modern processors usually put in their pipelines a dozen of instructions ahead of the executed one. The issue with branch like *if then else* is that the processors try to make a prediction on the branch that is going to be executed and start pipelining the predicted branch. If the prediction is right everything is fine, however if the prediction is wrong the whole pipeline might need to be flushed to execute the program correctly. This pipeline flush slows down the execution greatly

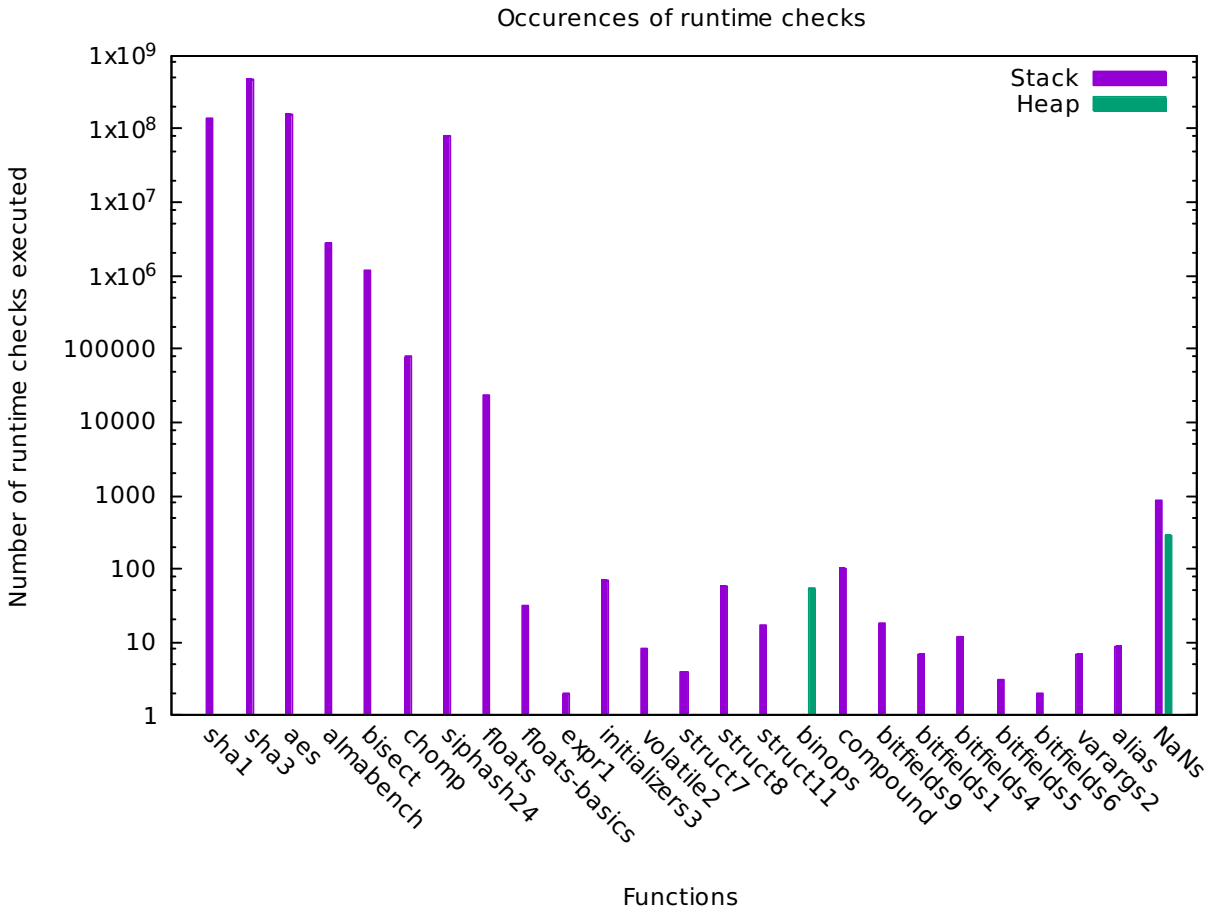


Figure 26: Number of runtime checks executed per program

which is why *if then else* are considered slow instructions. However in our implementations these predictions are certainly correct most of the time which would explain that branching introduces less overhead than the branchless version.

Relation between runtime checks and delay. The previous tests show that generally the implementations with runtime checks are slower than the one with fixed frames size without protection. We suspect that the number of runtime checks is related to the amount of delay introduced by our transformation. Especially the three programs cited earlier *sha1*, *sha3* and *aed* which have the most amount of runtime checks shows Figure 25 slower execution than "Fixed frames size".
To try to find a correlation we drew another graph Figure 27 representing the percentage of delay introduced in each program depending on the number of runtime checks added. Like before the number of runtime checks axis uses a logarithmic scale to include all obtained values.

The implementation chosen for the overhead is the protected one with *if then else* runtime checks since we concluded earlier that branchless are generally slower. For this graph we selected only programs that had a significant amount of execution time so the percentage of overhead calculated is not too biased due to the precision of our tool (*time* Unix utility). Furthermore among these

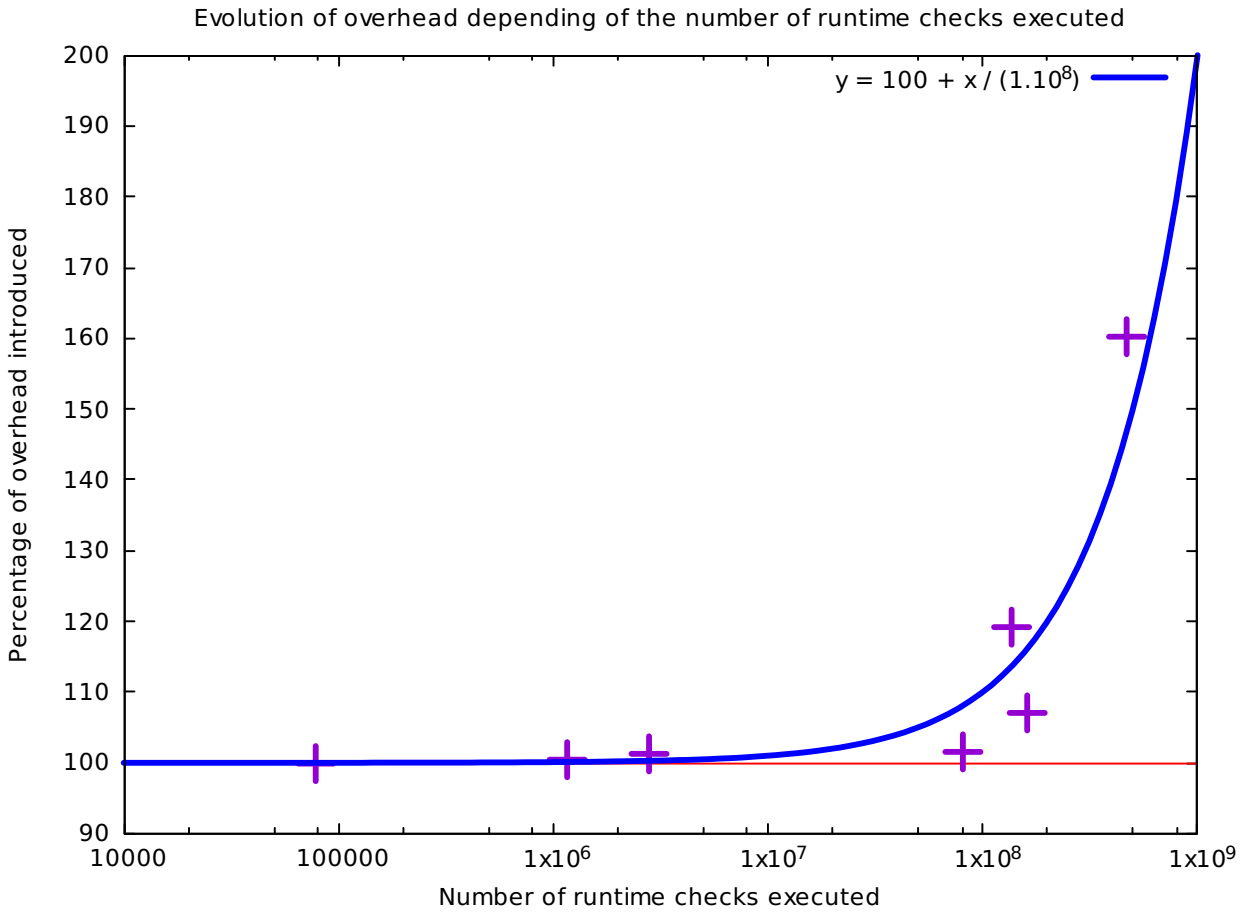


Figure 27: Evolution of the overhead percentage depending on the number of runtime checks executed

programs we only selected the one which had executed some runtime checks during their runtime. After the selection we were only left with seven programs to make a graph, the results are presented Figure 27.

We effectively see the percentage of overhead increasing with the number of runtime checks executed in our measures (purple cross). To find the correlation we try to find a relation which corresponds the most to the measures taken. The relation found is the blue curve and is represented by $y = 100 + x/(1.10^8)$. Even though we have a relatively low amount of points we can see that the blue curves matches their behaviour nicely. We have found a linear relation between the percentage of overhead and the number of runtime checks. Thus we can affirm that the overhead introduced is directly related to the number of runtime checks executed.

5.2.2 Stack frames size

Another figure we thought would be interesting was to see the distribution of the minimal frames size necessary for the programs to run. These data are represented Figure 28.

This gives us an idea about the size of the stack in the protected implementations. The results

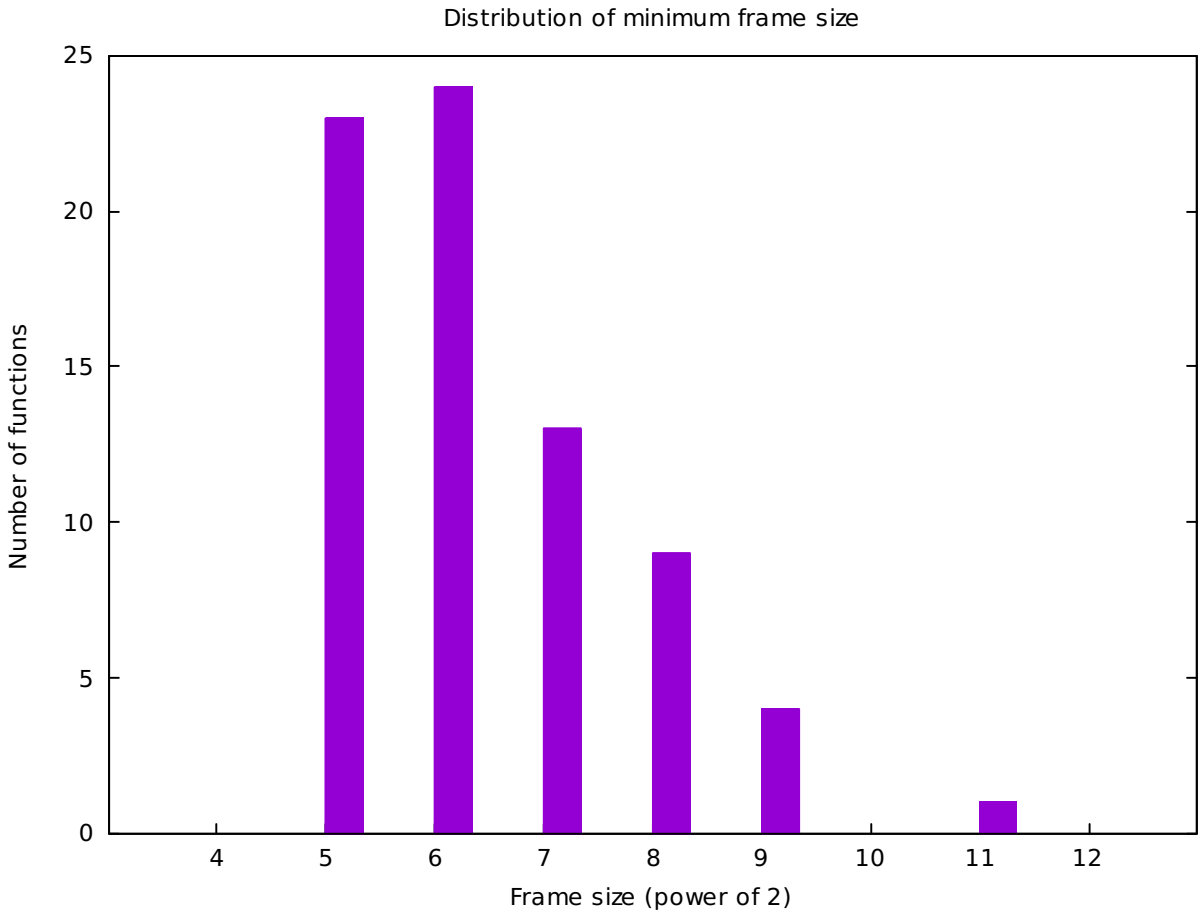


Figure 28: Distribution of the minimum frames size among 74 test programs

obtained show the minimal frames size required for our programs are small enough to not impact the memory usage. However the test programs we used were all relatively small and the distribution of bigger programs could be different. A more interesting graph related to memory consumption would be the comparison of the stack maximum growth between our implementations and the original CompCert.

5.3 Discussion

In this section we are going to talk about some points that our implementation is still lacking and also ways to improve our implementation that we could explore in the future.

More tests. Our implementation protects return addresses by modifying the stack layout of the programs. With this guarantee we are confident that we can prevent any kind of ROP attacks. However we only tested our implementation with simple buffer overflows. To show that our approach is correct, it would be interesting to pit our protection mechanisms against more complicated ROP attacks. We use CompCert so that eventually we will prove the security properties of our approach

. This would be an even more convincing way to affirm the benefits of our approach.

Guard area between the stack and the heap To ensure that either the stack or the heap grows too much and exceeds their designed area we want to put a guard area in the memory. The idea is to define a specific area in the memory, for example [`0xee000000 - 0xff000000`], where it is forbidden to write. If an instruction writes illegally in the guard area the program will detect it and will crash.
The principle is represented in Figure 29. Indeed we can see that the stack and the heap are clearly separated by the guard area located between [`0xee000000 - 0xff000000`]. Thus we are sure that every address above `0xff000000` is part of the stack.

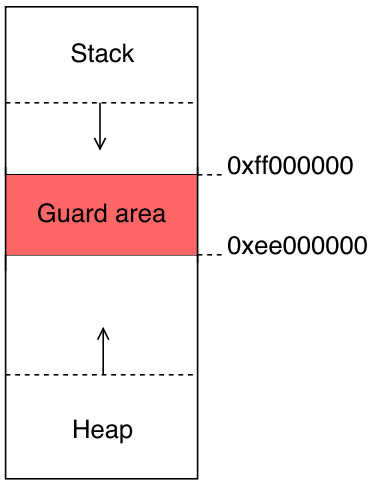


Figure 29: Guard area for the memory

To detect the write in the guard one possible way is to initialize the area with 0 for example. If we detect a bit in the guard area with the value 1 then we know that the guard has been corrupted and we make the program crash. Since our approach protects against ROP attacks, which takes effect when returning from a function, it is adequate to check the integrity of the guard area at the end of each function before returning. To be honest the guard area has not been implemented yet. Most of the programs do not have a stack or a heap which grows enough to exceed the limit of `0xff000000` so we were able to do satisfactory tests nevertheless. This guard area is necessary for our implementation to be complete and we hope that we will be able to do it during the remaining time of the internship.

Protection of runtime checks. Actually in our implementation we protect return addresses with runtime checks but we never protect these runtime checks. As in SFI we also need to prevent malicious instructions to avoid going through these checks. Therefore we need to make sure that is not possible for a program to freely jump in the code. This way an attacker would be able to bypass the security measures we established. In the C language the only expression that could allow such operations would be function pointers. Function pointers allow one to execute the code addressed by the pointer. The simplest solution would be to totally forbid the usage of function pointers which would solve our issue. We could also adapt ideas of SFI [5] like creating memory chunks and aligning them in order to render the runtime checks atomic. This idea is similar to our

approach, it would need to transform the structure of the memory for the code and then perform runtime checks to make sure we only jump to authorized locations.

Ways of improvement. We saw in the earlier section that our implementation was relatively slow compared to the original CompCert. Furthermore the amount of overhead is dependent of the number of runtime checks executed. The idea would be to either reduce the number of runtime checks effectively done or to improve the efficiency of our runtime checks. We can certainly narrow down the number of dangerous statements detected. For example Figure 30 presents some harmless C code that would be protected by our approach but does not present any vulnerability for a ROP attack. The pointer presented here points to an address of the heap so if the heap does not overflow on the stack it's not possible for the pointer *bar* to write on a return address as it is. We can see clearly that the protection mechanism that will be introduced in this code is not necessary and will just slow down the program. To detect these kind of occurrences, it would be necessary to make a static analysis of the code and pinpoint the statements that present a vulnerability for ROP attacks. If we could implement such static analyser find ref maybe the overhead from our implementation will be greatly reduced since we know the more we add runtime checks the more our program will slow down.

```
1 void foo(){
2         int *bar = malloc(sizeof(int));
3         *bar = 1;
4 }
```

Figure 30: Example of harmless C code

Another way is to improve the performance of our runtime checks. We tried to use a branchless implementation unsuccessfully. However another solution is to use a superoptimizer that can give us a super efficient assembly code equivalent to the C code used as input. If it finds a super effective assembly version of our runtime checks we could create a native function *runtime_check* in Cminor which will be expanded during the ASM compilation step.

Transparency. Our approach should leave safe programs behaviour unchanged. We do not want our implementation to change harmless programs in any way. Our runtime checks are transparent since the error behaviour is only triggered during illegal executions, otherwise the code executed will stay the same. The other modifications were done at the stack level, the stack does not exist at the level of the C language. Hence apart from inline assembly (that we want to forbid) it should not be possible to interact with the stack with C. The only way left would be a program which uses direct addressing, in other words a C program which gives raw addresses to pointers like the statement `pointer = 0xff000022`. Such program is considered unsafe anyway and with the existence of ASLR the behaviour of the program would be unpredictable.

For these reasons we are confident that our implementation is transparent, but a stronger guarantee would be to give a proof of transparency.

6 Conclusion

We presented in this paper a security technique called Software Fault Isolation which uses software methods to isolate the execution of dangerous modules. Hence SFI enables protected program to use external modules without taking the risk of compromising its execution. SFI has the benefits of demonstrating good performances concerning speed compared to other memory protection approaches. It also has a relatively small TCB. In this report we presented various implementations of SFI and especially the last one which used the certified compiler CompCert [3]. Even though SFI has many strong points it can be considered lacking when it comes to securing the return addresses. Indeed few solutions were explained to face this issue and those were usually not easily adaptable or resulted in lots of overheads. More importantly one the most common category of cyber attacks called ROP attacks explicitly target these return addresses to modify the flow control of the programs in their favour.

Therefore we propose a solution which is reliable and relatively efficient to protect return addresses from being corrupted. Our idea is to modify the structure of the stack in order to have an easy way to know the locations of all the return addresses. Afterwards we add runtime checks before the instructions that can compromise the integrity of the return addresses. These runtime checks take advantage of the new stack layout to forbid any illicit write on the return addresses locations.

We implementationed our approach with the compiler CompCert for the x86-32 architecture. Even though the implementation is not totally completed, the main work has been carried out and it successfully blocked programs containing simple buffer overflows. The different experimentations show that our transformations reduce the efficiency of our compiled modules depending on the number of runtime checks. Thus the decrease in efficiency varies greatly depending on the number of dangerous instructions that has to be tested in the compiled code.

However our implementation could still be improved by optimizing the runtime checks efficiency or by reducing the number of runtime checks added. For example, by using static analysis we could refine our detection of dangerous instructions and narrow down the number of runtime checks required which could greatly improve our performance.

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