An Obfuscating C Compiler for Encrypted Computing

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Abstract: This paper describes an 'obfuscating' C compiler for encrypted computing. The context consists of (i) a processor that 'works encrypted', taking in encrypted inputs and producing encrypted outputs while the data remains in encrypted form throughout processing, and (ii) machine codes that support arbitrary interpretations of the encrypted input and outputs from each instruction, as far as an adversary who does not know the encryption can tell. The compiler on each recompilation of the same source generates object code of the same form for which the runtime traces have the same form, but the data beneath the encryption may arbitrarily differ from nominal at each point in the trace, independently so far as the laws of computation allow.

1 INTRODUCTION

This article describes practical 'obfuscating' compilation from ANSI C (ISO/IEC, 2011) for encrypted computing (Breuer and Bowen, 2013). Encrypted computing generates security via hardware encryption for user data against powerful insiders such as the operator and operating system as adversaries. Its major component is a processor that 'works encrypted' in user mode and works unencrypted in the operator mode in which operating system code runs. Encryption keys are installed at manufacture, as with Smartcards (Körnerling and Kuhn, 1999), or uploaded in public view to a write-only internal store via a Diffie-Hellman circuit (Buer, 2006), and are not programmatically accessible. Theory in (Breuer et al., 2017) shows the operating system cannot directly or indirectly by deterministic or stochastic means read user data beneath the encryption, even to a probability slightly above chance. That turns out to imply it cannot be rewritten deliberately, even stochastically, to a value beneath the encryption that is independently defined, such as π, or the encryption key.

Several fast processors for encrypted computing are described in (Breuer et al., 2018), including the authors’ own KPU (Krypto Processor Unit) (Breuer et al., 2016), which runs at approximately the speed of a 433 MHz classic Pentium, embedding AES-128 (Daemen and Rijmen, 2002), and the slightly older HEROIC (Tsoutsos and Maniatakos, 2015) which runs like a 25 KHz Pentium, embedding Paillier-2048 (Paillier, 1999).

But the machine code instruction set defining the programming interface is also important. A conventional instruction set is insecure against powerful insiders who may steal an (encrypted) user datum x and put it through the machine’s division instruction to get \(x/x\) encrypted, an encrypted 1. Then any desired encrypted y may be constructed by repeatedly applying the machine’s addition instruction. By using the instruction set’s comparator instructions (testing \(2^{31} \leq z, 2^{30} \leq z, \ldots\) on an encrypted z and subtracting on branch, z may be obtained efficiently. That is a chosen instruction attack (CIA) (Rass and Schartner, 2016). The instruction set has to resist that, but the compiler must be involved too, else there would be known plaintext attacks (KPAs) (Biryukov, 2011) based on the idea that not only do instructions like \(x−x\) predictably favor one value over others (the result there is always \(x−x=0\)), but human programmers intrinsically use values like 0, 1 more often than others. The compiler’s job is to undo those statistics.

The compiler described here generates object code that varies from compilation to compilation of the same source code but always looks the same to an adversary, the difference lying in encrypted constants that the adversary cannot read. Runtime traces also ‘look the same,’ with the same instructions (modulo encrypted constants) in the same order, the same jumps and branches, reading from and writing to the same registers. But data beneath the encryption varies arbitrarily and independently from compilation to compilation at each point in the trace, subject only to the proviso that a copy instruction preserves value, and the variation at the start and end of a loop is equal.

The compiler does that even for object code such as the single instruction that adds two numbers, compiled from source code \(x = y + z\). That gives the following conditions on the machine code instruction design described in (Breuer et al., 2017) and Box 1: in-
instructions must execute atomically (1), or recent attacks such as Meltdown (Lipp et al., 2018) and Spectre (Kocher et al., 2018) against Intel might become feasible, must work with encrypted values (2) or an adversary could read them, and must (3) be adjustable via embedded encrypted constants to offset the values beneath the encryption arbitrarily. The condition (4) is for the security proofs in (Breuer et al., 2017), and amounts to different padding or blinding factors for encrypted program constants and runtime values.

The effect of (1-4) is that an adversary not privy to the encryption can feasibly believe nearly anything of the data values beneath the encryption in a runtime trace ((*) of Box 2). By (3), an arbitrary variation from nominal introduced in one instruction could be corrected and changed again in the next instruction, and (1-2) prevent the adversary noticing directly. The compiler’s job then boils down to varying the encrypted program constants from recompilation to recompilation so that all the feasible variations in the runtime data are not only achieved in some recompilation but also equiprobable across recompilations ((§) of Box 2). It is shown in (Breuer et al., 2017) that implies cryptographic semantic security (Goldwasser and Micali, 1982) for user data against insiders not privy to the encryption (fix) of Box 2). I.e., encrypted computation does not compromise the encryption.

How the compiler achieves the ‘equiprobable variation’ required by (‡) is encapsulated in Box 3.

It generates a new obfuscation scheme for the object code each recompilation. That is an offset for the data beneath the encryption in every memory and register location per every point in the program control graph.

Specifically, the compiler C[–] translates an expression e that is to end up in register r at runtime into machine code mc and generates a 32-bit offset ∆e for r at the point in the program where it is loaded with the result of the expression e. That is

\[ C[e] = (mc, ∆e) \]  

The offset ∆e is the part of the obfuscation scheme relating to register r at the point where the encrypted value of the expression is written to it.

Let \( s(r) \) be the content of register r in state s of the processor at runtime. The machine code \( mc \)’s action changes state \( s_0 \) to an \( s_1 \) with a ciphertext in \( r \) whose plaintext value differs by \( ∆e \) from the nominal value \( e \) (bitwise exclusive-or or the operator of another mathematical group are alternatives to addition here):

\[ s_0 \xrightarrow{mc} s_1 \text{ where } s_1(r) = E[e + ∆e] \]  

The encryption \( E \) is shared with the user and the processor (but not the potential adversary, the operator and operating system). The randomly generated offsets \( ∆e \) of the obfuscation scheme are known to the user, but not the processor and not the operator and operating system. The user compiles the program and sends it to the processor to be executed and needs to know the offsets on the inputs and outputs. That allows the right inputs to be created and sent off for processing on the encrypted computing platform, and allows sense to be made of the outputs received back.

### 1.1 Article Organisation

Section 2 introduces a modified OpenRISC (http://openrisc.io) machine code instruction set for encrypted computing satisfying (1-4) first described in (Breuer et al., 2017), and its abstract semantics. That will be used instead of assembly language to represent machine code, so readers need not face assembler.

What is the difficulty solved here? It is that nobody knows how to compile satisfying (A-C) of Box 3, or even if it is possible. (Breuer et al., 2017) showed how to do it for structured ‘if’ and ‘while’ statements in C, assignments to variables, sequences
of statements, and function calls and return. The only plaintext type treated was the 32-bit signed integer type (‘int’). But extension to such C features as pointers (memory addresses) is difficult, since an obfuscating offset in a memory address amounts to a miss, where an offset in the value in that location is correctable.

Section 3 resumes obfuscating compilation up to (Breuer et al., 2017), then extends it to nearly all ANSI C, including pointers and arrays, union and ‘struct’ (record) types, floats, double length integers and floats, short integers and chars, both signed and unsigned, gotos, and the interior functions extension.

1.2 Notation

Encryption is denoted by $x' = E[x]$ of plaintext value $x$, sometimes just $x'$. Decryption is $x = D[x']$. The operation on the ciphertext domain corresponding to $f$ on the plaintext domain is written $[f]$, where $x'[f]y' = E[x f y]$. Given register $r$, $r'$ is the next in sequence. Encrypted offsets $E[dx]$ may be compactly written $Δ'$.  

2 FXA INSTRUCTIONS

A ‘fused anything and add’ (FxA) (Breuer et al., 2017) ISA is the compilation target here, satisfying conditions (1-4) of Section 1. Most of the integer portion is shown in Table 1. It is adapted from the OpenRISC instruction set v1.1 https://openrisc.io/orlk.html. That has about 200 instructions (6-bit opcode plus variable minor opcodes) separated into single and double precision integer and floating point and vector subsets and the instructions are all 32 bits long. Instructions access up to three 32 general purpose registers (GPRs), defined via 5-bit specifier fields. A register arithmetic such as addition $x ← y + z$ will specify three GPRs in which $x$, $y$, $z$ reside in $3 \times 5 = 15$ bits, while the operation ‘+’ is specified in the 6-bit primary opcode plus an 11-bit minor. The minor may specify if addition is signed or unsigned, half, single or double precision, integer or floating point, for example. One operand may be supplied as a (‘immediate’) constant in the instruction itself.

FxA instructions may contain 128-bit or more encrypted constants, so some shoehorning is required. A ‘prefix’ instruction takes care of that, supplying extra bits as necessary. The prefix instruction is 32 bits long, but several may be concatenated.

In addition to the integer instructions of Table 1, there are floating point instructions addf, subf, mulf etc. paralleling the OpenRISC floating point subset. Contents of registers and memory are encryptions of 32-bit integers that encode floating point numbers (21 mantissa bits, 10 exponent bits, 1 sign bit) via the IEEE 754 floats, and let $xR = E[x]$, $R$ encrypted value $E[R]$, $x[R]$ $x$ $xR = x R$.

### Table 1: Integer portion of FxA machine code instruction set for encrypted working – abstract syntax and semantics.

| op. fields | mem. semantics |
|------------|----------------|
| add $r_0, r_1, r_2, r'_0$ | $r_0 ← r_1 + r_2 + r'_0$ |
| sub $r_0, r_1, r_2, r'_0$ | $r_0 ← r_1 - r_2 - r'_0$ |
| mul $r_0, r_1, r_2, r'_0$ | $r_0 ← (r_1 \cdot r_2) + r'_0$ |
| div $r_0, r_1, r_2, r'_0$ | $r_0 ← (r_1 \div r_2) + r'_0$ |
| beq $r_0, r_1, r_2, r'_0$ | if $b$ then $r_0 ← r_1$ if $b$ then $r_0 ← r_2$ |
| bne $r_0, r_1, r_2, r'_0$ | if $b$ then $r_0 ← r_1$ if $b$ then $r_0 ← r_2$ |
| mov $r_0, r_1$ | move $r_0 ← r_1$ |
| beq $r_0, r_1, r_2, r'_0$ | branch if $b$ then $r_0 ← r_1$ |
| bne $r_0, r_1, r_2, r'_0$ | branch if $b$ then $r_0 ← r_2$ |
| j $r_0, r_1, r_2, r'_0$ | jump $r_0 ← r_1, r_2$ |
| jr $r_0, r_1, r_2, r'_0$ | jump $r_0 ← r_1, r_2$ |
| b $j, r_0, r_1, r_2, r'_0$ | branch $pc ← pc + j$ |
| sw $(r_0, r_1, r_2, r'_0)$ | store $mem[r_2] ← r_1$ |
| lw $(r_0, r_1, r_2, r'_0)$ | load $r_0 ← mem[r_1] + r_2$ |
| jr $r_0, r_1, r_2, r'_0$ | jump $pc ← pc + r$ |
| jal $r_0, r_1, r_2, r'_0$ | jump $ra ← pc + 4$, $pc ← j$ |
| j $j, r_0, r_1, r_2, r'_0$ | jump $pc ← j$ |
| nop $r_0, r_1, r_2, r'_0$ | no-op |

### Legend

- $r$ – register indices
- $k$ – 32-bit integers
- $pc$ – prog. count reg.
- $j$ – program count or incr.
- $ra$ – return addr. reg.
- $E[.]$ – encryption
- $D[.]$ – decryption
- $Δ'$ – register content
- $x[R]$ – encrypted value $E[R]$, $x[R]$ $x$ $xR$ – $xR = x R$
and first offsets them as integers, then multiplies them as floats, finally offsetting as integer again. The operation is atomic, leaving no trace if aborted.

Our own FxA set has an improvement for runtime efficiency. Floating point comparisons (indeed, also unsigned integer comparisons) have two encrypted constants (the signed integer comparisons in Table 1 only have one). The floating point branch-if-equal instruction, for example, internally calculates ($\dagger = f$):

$$ (r_1 - k_1) \land (r_2 - k_2) \quad (\dagger = f)$$

where $= f$ is the floating point comparison on integers encoding floats via IEEE 754, and $\land = f$ is the corresponding test in the ciphertext domain. That is, $x'[\land = f]' \iff x = f y$.

Condition (2) of Section 1 implies there should be one more constant, an encrypted bit $k_0'$ that decides if the 1-bit result of the test is to be inverted. That is because the test outcome is observable by whether the branch is taken, so by condition (3) of Section 1 it should be varied by an encrypted constant in the instruction. Our own FxA instruction set does not have that constant, because the variation is taken care of by the compiler. It independently and randomly decides each time if the underlying source code condition is to be interpreted as though by a truth teller, who says ‘true’ when true is meant and ‘false’ when false is meant, or by a liar, who says ‘false’ when true is meant and ‘true’ when false is meant. It equiprobably generates code for a lie and uses the branch-if-not-equal instruction, or truthful code and uses the branch-if-equal instruction. The compile procedure is described exactly in (Breuer et al., 2017). With this arrangement, whether the branch happens or not at runtime does not even statistically relate to what the boolean result is meant to be. Then condition (3) of Section 1 on the ‘output’ of the instruction being encrypted is satisfied vacuously, as there is effectively no output. An observer who sees a branch does not know if that is the result of truthful interpretation of the source code and the condition has come out true at runtime, or it is the result of mendacious interpretation and the condition has come out false at runtime.

A different FxA implementation may prefer to encode an extra bit in the constants $k_1$, $k_2$, the encrypted versions $k_1'$, $k_2'$ of which appear in the instruction, and use the exclusive or of the two to determine if equals or not-equals is executed, satisfying (3) literally.

3 OBfuscating compilation

The compiler works with a database $D : \text{Loc} \rightarrow \text{Int}$ containing the (32-bit) integer (type Int) offsets $\Delta_l$ for data, indexed per register or memory location $l$ (type Loc). The offset represents by how much the runtime data underneath the encryption is to vary from nominal at that point in the program, and the database $D$ comprises the obfuscation scheme. It is varied randomly by the compiler as it makes its pass. The compiler also maintains a database $L : \text{Var} \rightarrow \text{Loc}$ binding source variables in registers and memory locations.

In (Breuer et al., 2017), a (non-side-effecting) expression compiler that places its results in register $r$ is set out, of type:

$$ C^L : \text{DB} \times \text{Expr} \rightarrow \text{MC} \times \text{Int} \quad (7)$$

where $\text{MC}$ is the type of machine code, a sequence of FxA instructions $mc$, and $\text{Int}$ contains the plaintext integer offset $\Delta$ from nominal that the compiler intends will exist in the result in $r$ beneath the encryption when the machine code is evaluated at runtime. The aim is to satisfy (§) by varying the offset $\Delta$ arbitrarily and equiprobably.

To translate $x + y$, for example, where both $x$ and $y$ are signed integer expressions, the compiler first emits machine code $mc_y^r$ computing expression $x$ in register $r$ with offset $\Delta x$. It then increments the register index $r$ to $r'$ (the processor contains registers allocated for this kind of temporary workspace calculation) and emits machine code $mc_y'^r$ computing expression $y$ in register $r'$ with offset $\Delta y$. That is

$$ (mc_y^r, \Delta x) = C^L[D : x]^r$$
$$ (mc_y'^r, \Delta y) = C^L[D : y]'^r$$

It then randomly decides an offset $\Delta$ for the whole expression $x + y$ and emits the FxA integer addition instruction (the abstract semantics $r \leftarrow r' [+][+][+]k'$ is presented here; notation as in Section 1.2) on the registers $r$ and $r'$ to return the result in $r$:

$$ C^L[D : x + y]^r = (mc, \Delta) \quad (8)$$
$$ mc = mc_y^r; mc_y'^r ; r \leftarrow r' [+][+][+]k'$$
$$ k = \Delta - \Delta x - \Delta y$$

If the expression were instead $x \times y$, then the emitted code terminates with a multiplication instruction $c$:

$$ C^L[D : x \times y]^r = (mc, \Delta) \quad (9)$$
$$ mc = mc_y^r; mc_y'^r ; c$$
$$ c = r \leftarrow (r [-][+]\Delta_y^l)'[+]r' [-][+]\Delta_y^l)'[+]\Delta'$$

In every case the final offset $\Delta$ for the runtime result in $r$ beneath the encryption is freely generated (§).
When an expression is a source code variable \( x \) on its own, the lookup database \( L \) serves to locate where the variable is currently held in registers and memory and the compiler will emit code to bring it into register \( r \) via an addition that introduces a new offset \( \Delta \), taking into account the offset \( \Delta' = DI \) for the value in storage in location \( l = Lz \). A compiler abstraction layer called RALPH (Register Abstraction Layer for Physical Hardware) optimises the positioning of data and also provides an effectively unlimited supply of temporary register space by aliasing-in use of the 64K extra ‘special purpose registers’ (SPRs) specified in the OpenRISC v1.1 architecture as required. If those are exhausted, it will interpose use of the execution stack (that is in memory). The compilation of the subexpressions \( x, y \) may use RALPH registers \( r, r', r'' \) respectively, but not ‘lower’ indices, so the code generated for \( x \) in \( r \) is not interfered with at runtime by the code generated for \( y \) in \( r' \).

The general rule is that the emitted code uses each opportunity of a write instruction to generate a new offset \( \Delta \) in the written location, fulfilling (§).

Statements are compiled differently to expressions because they do not produce a result for which an offset from nominal is to be generated. Instead they have a side-effect. Let Stat be the type of statements, then compiling a statement returns not a single offset, as for expressions, but a new obfuscation scheme:

\[
\text{C}^L[\cdot : \cdot] : \text{DB} \times \text{Stat} \rightarrow \text{DB} \times \text{MC}
\]

Recall that the database \( D \) of type DB holds the obfuscation scheme (the offsets from nominal values in all locations, per register or memory location).

Consider an assignment \( x = e \) to a source code variable \( x \), which the location database \( L \) says is bound in register \( r = Lx \). A pair in the cross product output will be written \( D : m \) as syntactic sugar. First code \( mc_e \) for evaluating expression \( e \) in temporary register \( t0 \) at runtime is emitted via the expression compiler:

\[
(mc_e, \Delta e) = C^L[D_0 : e]^{10}
\]

The offset \( \Delta e \) is generated by the compiler for the result \( e \) in \( t0 \). An instruction to correct it to a new randomly chosen offset \( \Delta \) in register \( r \) is emitted next:

\[
\text{C}^L[D_0 : x = e] = D_1 : mc_e; r \leftarrow t0 \,[+] k'\]

The change made in the database of offsets \( D_0 \) to \( D_1 \) is at the entry for location \( r \), where the initial offset \( D_{0r} \) changes to \( D_{1r} = \Delta \), the new offset being freely and randomly chosen by the compiler, supporting (§).

### 3.1 Other Base Types

Double length (64-bit) ‘long long’ integers are encoded as two 32-bit integers. The FxA instructions for dealing with them contain (encrypted) 64-bit constants. Encryption \( x' = E[x] \) is extended to 64-bit integers as two encrypted 32-bit integers \( x' = E[x^H]; E[x^L] \) where \( x^H, x^L \) are the high and low 32 bits of the 64-bit plaintext integer \( x \) respectively.

Let \(- and + be the usual subtraction and addition on 32-bit integers, and let \(-^2 and +^2 be their application two-by-two to the pairs of 32-bit integers comprising 64-bit integers. Let \( r_H \) be the usual multiplication on the 64-bit ‘long long’ integers. The FxA 64-bit multiplication instruction has semantics:

\[
x' \leftarrow E[(y - z)k_1] \times (z - k_2) + k_0 \quad (\ast y)\]

where \( k_0, k_1, k_2 \) are 64-bit plaintext integer constants with the encrypted ciphertexts \( k'_i \), \( i = 0, 1, 2 \) embedded in the instruction.

Putting it in terms of the effect on register contents, the instruction semantics is:

\[
r_H^2 \cdot r_0 \leftarrow (r_H^2 \cdot r'_0 \times [-2] k'_1) \times r_2 \times [2] k'_2 + [2] k'_0
\]

However, for (encrypted) 64-bit operations the processor partitions the register set into twos referred to by one name each. In those terms the semantics is:

\[
r_0 \leftarrow (r_1 \times [-2] k'_1) \times r_2 \times [-2] k'_2 + [2] k'_0
\]

When emitting the instruction, the compiler understands that the registers \( r \) listed are really the firsts of pairs \( r, r' \) containing the (encrypted) 64-bit integer as two (encrypted) 32-bit halves, the high portion in \( r \) and the low portion in \( r' \) (bigendian order). So for 64-bit multiplication the compiler emits not the instruction \( e \) of (9) in which successive registers \( r \) and \( r' \) are used for subexpressions \( x \) and \( y \) respectively, but one in which the registers \( r \) and \( r'' \) are used, respectively:

\[
e = r \times \times [-2] k'_1 \times r_2 \times [-2] k'_2 \times [2] k'_0
\]

That leaves enough space for \( x \) in \( r \), \( r'' \) and \( y \) in \( r'' \), \( r'' \). The offset \( \Delta \) is chosen freely, satisfying (§). As already set out for \( \times \) and \( \times \), single precision floats are encoded as 32-bit integers according to the IEEE 754 standard. Double precision 64-bit floats (‘double’) are encoded as two (encrypted) 32-bit integers, the top and bottom halves respectively of a 64-bit integer following the IEEE 754 coding for the double. Let \( \times_d \) operation be the double precision float multiplication expressed on the 64-bit integer encodings of doubles as two 32-bit integers, and let \( \times_d \)
be the corresponding operation in the cipherspace domain on two pairs of encrypted 32-bit integers. Then as with double-length integer multiplication (12), the compiler emits no instruction $c$ of (9) with registers $r$ and $r'$ for subexpressions $x$ and $y$ respectively, but

$$c = r \leftarrow (r[-\Delta']_s) [s] d (r[r'[\Delta']_s] [\Delta']_s) \leftarrow (r[r'[\Delta']_s] [\Delta']_s)$$ (13)

with $r$, $r'$, leaving room for double length operands.

Machine code instructions that act on ‘short’ (16-bit) or ‘char’ (8-bit) integers are not needed because short integers are promoted to the standard 32-bit length at each operation in C. The compiler has a repertoire of translations for the source code cast operation that change the 13 basic C types (signed/unsigned char, short, int, long, long long integer, and float and double precision float, also the single bit bool type) into one another. A short is a (encrypted) 32-bit integer with the top 16 bits ignored. To cast a signed int to a signed short, for example, the compiler in principle issues code that moves the integer 16 places left and then 16 places right again:

$$C^L[D : (short)e]' = (mc, \Delta)$$

$$C^L[D : e]' = (mc, \Delta)$$

Then pointer-based access becomes easier to generate code for. At compile time where in the array the pointer will point is unknown, but the single shared offset may be used. The downside is that pointers $p$ must be declared with an array into which they point:

restrict $A$ int *$p$;

With this approach, the compiler takes into account the offset for $A$ when constructing the dereference $*e$ of an expression $e$ that is a pointer into $A$ as follows. It first emits code $mc_e$ that evaluates the pointer with a randomly generated offset $\Delta_e$ beneath the encryption:

$$(mc_e, \Delta_e) = C^L[D : e]'$$

Then it emits a load instruction that compensates for the offset $\Delta_e$ in the address. The load instruction semantics is overall as follows— but as the text below explains, the runtime evaluation is not simple-minded:

$$r \leftarrow (r[-\Delta'_s]) [-2] [16] [+2]\Delta'_s$$

The constants $k$, $\Delta$ are freely chosen.

In order to avoid encodings of numbers like $2^{16}$ appearing more than averagely often, the compiler in fact rather than emit the literal sequence above causes a register $r'$ to be loaded with the encryption of a random number $k_1$ and uses the instructions with $r'[-\Delta'_s]$ in place of $E[2^{16}]$ where $k_2=k_1-2^{16}$. That makes the encrypted constants that appear embedded in machine code unbiasedly distributed (recall that (4) says they cannot serve as data in the processor arithmetic).

For casts between integer and floating point types, the instruction set provides atomic integer-to-float (and vice versa) conversions. In accord with the principle (3) of FxA design, those embed encrypted constants that displace inputs and outputs.

### 3.2 Arrays and Pointers

It is possible to represent an array $A$ as a set of individual variables $A_0, A_1, \ldots$ and that allows the compiler to translate a lookup $A[i]$ as a compound expression $'i='0)'A0':(i=')A1:\ldots'$. While a write $A[i]=x$ can be translated to $'i='0)'A0:='x'else if (i='1)$ then $A1='x$ else $\ldots'$. The entries get individual offsets from nominal $\Delta A_0, \Delta A_1, \ldots$ in the obfuscation scheme maintained by the compiler.

While that is logically correct, array access should be better than $O_n$, so we have explored a different approach: entries of array $A$ all share the same offset $\Delta A$ from their nominal value (beneath the encryption).

Then pointer-based access becomes easier to generate code for. At compile time where in the array the pointer will point is unknown, but the single shared offset may be used. The downside is that pointers $p$ must be declared with an array into which they point:

restrict $A$ int *$p$;

With this approach, the compiler takes into account the offset for $A$ when constructing the dereference $*e$ of an expression $e$ that is a pointer into $A$ as follows. It first emits code $mc_e$ that evaluates the pointer with a randomly generated offset $\Delta_e$ beneath the encryption:

$$(mc_e, \Delta_e) = C^L[D : e]'$$

Then it emits a load instruction that compensates for the offset $\Delta_e$ in the address. The load instruction semantics is overall as follows—but as the text below explains, the runtime evaluation is not simple-minded:

$$r \leftarrow (r[-\Delta'_s]) [-2] [16] [+2]\Delta'_s$$

The processor at runtime has the encrypted address $a'$ of the intended entry at address $a$ of array $A$.

$$a' = r[-\Delta'_e]$$

Despite encryption being one-to-many, in the case of the KPU at least, the processor’s internal access to keys makes that into a unique value $\hat{a}$ (not necessarily the decrypted address $a$, though it is a unique hash of that in the KPU in case of symmetric encryption; for Paillier it is a partial decryption that removes an extra ‘blinding’ multiplier from the ciphertext).

In the KPU, the value $\hat{a}$ is memoised within a special front end to the address translation lookaside buffer (TLB) within the processor to a value $\text{TLB}(\hat{a})$ selected at first encounter. It was merely the next location free in a contiguous memory area in RAM, so there is no mathematical relation of $a$ with $\text{TLB}(\hat{a})$, which RAM gets as the address to look up. The load instruction exact semantics is then as follows:

$$r \leftarrow \text{mem}[\text{TLB}(\hat{a})]$$

The TLB capacity in the KPU is 2M addresses. It is backed by an encrypted database in RAM, and a
missing memoisation causes a minor memory fault signal, the handler for which recovers the encrypted information from RAM to the TLB.

The memoisation is changed at every write to it, so RAM sees a random pattern of accesses overall, but here for read the existing memoisation TLB is used. The upshot is that the simple load instruction issued by the compiler both works at runtime despite RAM getting a dynamically varying address to lookup, and the location containing encrypted data, and also does not expose unencrypted information.

The value retrieved by the load instruction has the offset $\Delta A$ and the compiler emits an add instruction to change it to a freely chosen offset $\Delta$, as follows:

$$r \leftarrow r[+]k'$$

where $k = \Delta - \Delta A$. The complete code emitted is:

$$C^L[D : \text{e}]^r = (mc, \Delta)$$

$$mc = mc_e;$$

$$r \leftarrow \text{mem}[r[-]A'] + \sigma \# \text{lw} r (E[-A_e])r$$

$$r \leftarrow r[+]k';$$

$$\# \text{addi} r r k'$$

where $k=\Delta-\Delta A$.

Then an indexed array lookup $A[i]$ is handled by dereferencing a pointer *(A+i). However, writing an array entry is more problematic, not because the code is complicated (the final addition instruction in (15) is replaced with an addition instruction involving the register $r'$ in which the overwriting value has been written), but because it should on principle be coincident with a change of the offset $\Delta A$ for the target entry in the array, and therefore for every entry in the array. That means every array entry must be rewritten to the new offset whenever one is written, an On ‘write storm.’ We are evaluating the real costs – only write bandwidth is occupied, not write latency increased, since writes are asynchronous through the writeback cache. The behaviour is like oblivious RAM (ORAM) (Ostrovsky, 1990), a modified RAM that encrypts data and addressing, hiding programmed accesses among random false accesses.

### 3.3 Structs

C ‘structs’ are records with fixed fields. The approach the compiler takes is to maintain a different offset per field, per variable of struct type. That is, for a variable $x$ of struct type with fields $a$ and $b$ the compiler maintains offsets $\Delta A_a$ and $\Delta A_b$. It is as though there were two variables, $x.a$ and $x.b$ respectively.

In the case of an array $A$ the entries of which are structs with fields $a$ and $b$, the compiler maintains two separate offsets $\Delta A_a$ and $\Delta A_b$ for the two fields of its entries, and so on recursively if those fields are themselves structs. Updating one field in one entry changes the offset and is accompanied by a ‘write storm’ of adjustments over the stripe through the array consisting of that same field in all entries. That is more efficient than a storm to a whole array, so for more efficient computing in this context, array entries should be split into structs whenever possible.

### 3.4 Unions

The obfuscation scheme in a union type such as

```
union { struct{int a; float b[2];}; double c[2]; }
```

has to engage compatible obfuscation schemes for the component types. The obfuscation scheme for the struct will have the pattern (in 32-bit words) $x,y,y$, with $x$ the offset for the int and $y$ the offset for the float array entries, while the pattern for the double array will be $u,v,u,v$, repeating $u,v$ for each entry.

```
union{struct{ int a;float b[2]; };double c[2]; }
```

The resolution is $x\leftarrow y\leftarrow u\leftarrow v$ for a scheme $y,y,y,y$. That is the least restrictive obfuscation scheme forced by the union layout here.

### 3.5 Go To and Come To

Gotos are treated in three steps: (i) local labels $x$ are declared as follows:

```
  __label__ x
```

Variables $y$ that might be affected (are read or written) after the labelled point ‘$x:...$’ in the program by a goto $x$ must be declared prior to the `__label__ x` declaration, because the compiler snapshots the obfuscation scheme (the offsets $\Delta y$) for the variables $y$ that are in scope at that point and binds it to $x$. Then (ii) at each goto $x$ statement the compiler emits machine code add instructions that change the offsets $\Delta y$ back to the scheme bound to $x$. Further (iii), just before label $x$ on the default ‘fall through’ control path, the compiler emits machine code add instructions that changes the offsets $\Delta y$ of the variables $y$ back to the scheme bound to $x$. So a label is associated with a non-trivial sequence of machine code, and it may be thought of as a source-level ‘come to $x$’ statement.

The reason for this approach is that converging control paths must agree on the obfuscation scheme
at the join. At the point labelled by x, paths via incoming gotos converge with the fall-through. What obfuscation scheme should be used there? The one at the goto x or the one at the labelled point? Either order would sometimes force two passes. The simplest approach is the one adopted: declare the label and bind the obfuscation scheme in use at the declaration.

3.6 Interior Functions

Interior functions are a common extension to ANSI C. They are functions f declared within another function g. They have access to g’s arguments and local variables x. At the point of declaration of f the compiler knows the obfuscation scheme $\Delta_x$ in use for g’s variables x and can emit code for the interior function f that is accommodated to it. Unfortunately, g may call f at a point where the obfuscation scheme for x is different, having been altered by the compiler at some intervening write to x in g.

The same solution as for gotos works. The interior function declaration is treated as a label declaration:

```
_LABEL_ f
```

The obfuscation scheme at that point is bound to f, and every call of f in g is treated as though it were preceded by the ‘come to f’ statement of Section 3.5. That reestablishes the obfuscation scheme bound to f at its declaration. A return from f is accompanied by instructions that establish another obfuscation scheme for the external’s variables x.

4 SUMMARY

Every machine code write instruction emitted by the compiler establishes a new offset $\Delta$ from the nominal value of the plaintext data beneath the encryption in the target location at that point in the program. The $\Delta$s are randomly chosen by the compiler, subject to the constraint that the offsets at the ends of control paths agree at the join (§). A function call to f does establish the same obfuscation scheme after every call, but different instances $f_i$ of f may be compiled per call point i, each with separate final obfuscation schemes.

At the current stage of development, the compiler has near total coverage of ANSI C and GNU extensions, including statements-as-expressions and expressions-as-statements. Efficient strings (currently arrays of chars) are still to come. It is being debugged via the gcc ‘c-torture’ testsuite v2.3.3, and we are presently about one quarter through that.

5 CONCLUSION

How to compile C in an ‘obfuscating’ way for encrypted computing has been set out. That class of platform natively works on encrypted data in user mode, never decrypting it as far as can be told via the programming interface. With a modified RISC instruction set for encrypted computing, arbitrary assignments for the plaintext data beneath the encryption in a single execution are feasible. The obfuscating compiler makes them all equiprobable. Without it, stochastic arguments are unavailable. With it, semantic security (no attack does better than guesswork) for user data beneath the encryption holds by definition, given the encryption is independently secure.

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