Borrowing Safe Pointers from Rust in SPARK

Georges-Axel Jaloyan¹*, Yannick Moy², and Andrei Paskevich³,⁴

¹ École Normale Supérieure,  
PSL Research University, Paris, France  
² AdaCore, Paris, France  
³ Laboratoire de Recherche en Informatique,  
Université Paris-Sud, CNRS, Orsay, F-91405  
⁴ Inria Saclay, Université Paris Saclay, Palaiseau, F-91120

Abstract. In the field of deductive software verification, programs with pointers present a major challenge due to pointer aliasing. In this paper, we introduce pointers to SPARK, a well-defined subset of the Ada language, intended for formal verification of mission-critical software. Our solution uses a permission-based static alias analysis method inspired by Rust’s borrow-checker and affine types, and enforces the Concurrent Read, Exclusive Write policy. This analysis has been implemented in the GNAT Ada compiler and tested against a number of challenging examples. In the paper, we give a formal presentation of the analysis rules for a miniature version of SPARK and prove their soundness. We discuss the implementation and compare our solution with Rust.

1 Introduction

SPARK [1] is a subset of the Ada programming language targeted at safety- and security-critical applications. SPARK restrictions ensure that the behavior of a SPARK program is unambiguously defined, and simple enough that formal verification tools can perform an automatic diagnosis of conformance between a program specification and its implementation.

As a consequence of SPARK’s focus on automation and usability, it forbids the use of programming language features that either prevent automatic proof, or make it possible only at the expense of extensive user effort in annotating the program. The lack of support for pointers in SPARK is the main example of this choice. While it is possible to exclude parts of the programs that manipulate pointers from analysis, it would be preferable to support pointers when their use does not prevent formal verification.

Among the various problems related to the use of pointers in the context of formal program verification, the most difficult problem is the possibility that two names refer to overlapping memory locations, a.k.a. aliasing. Formal verification platforms that support pointer aliasing like Frama-C [2] require users to annotate

* This work is partly supported by the Joint Laboratory ProofInUse (ANR-13-LAB3-0007) and project VECOLIB (ANR-14-CE28-0018) of the French National Research Agency (ANR).
programs to specify when pointers are not aliased. This can take the form of inequalities between pointers when a typed memory model is used, or the form of separation predicates between memory zones when an untyped memory model is used. In both cases, the annotation burden is acceptable for leaf functions which manipulate single-level pointers, and quickly becomes overwhelming for functions that manipulate pointer-rich data structures. In parallel to the increased cost of annotations, the benefits of automation decrease, as automatic provers have difficulties reasoning explicitly with these inequalities and separation predicates.

Programs often rely on non-aliasing in general for correctness, when such aliasing would introduce interferences between two unrelated names. We call such aliasing potentially harmful when a memory location modified through one name could be read through another name, within the scope of a verification condition. Otherwise, the aliasing is benign, when the memory location is only read through both names. A reasonable restriction for formal program verification is thus to forbid potentially harmful aliasing of names. The difficulty is then to guarantee the absence of potentially harmful aliasing. The following code shows an example where we want analysis to be able to rely on the non-aliasing of parameters \( X \) and \( Y \) to prove the postcondition of the procedure \texttt{Assign_Incr}:

```plaintext
procedure Assign_Incr (X, Y : in out Integer_Pointer )
    with Post => Y.\text{all} = X.\text{all} + 1
is
begin
    Y.\text{all} := X.\text{all} + 1;
end Assign_Incr;
```

In this work, we present the first step for the inclusion of pointers in the Ada language subset supported in SPARK. As our main contribution, we show that it is possible to borrow and adapt the ideas underlying the safe support for pointers in permission-based languages like Rust, to safely restrict the use of pointers in usual imperative languages like Ada. This adaptation is based on a possible division of work between a permission-based anti-aliasing analysis, lifetime management by typing, and the use of a formal verification platform for checking non-nullity of accessed pointers. For example, these rules prevent aliasing between parameters \( X \) and \( Y \) in the code of procedure \texttt{Assign_Incr} above, which makes it possible to treat pointers in proof like records with a single field corresponding to the type of the object pointed to. Thus, the verification condition corresponding to the postcondition of procedure \texttt{Assign_Incr} has a form (using \texttt{get/set} to access the field \texttt{all} of variables \( X \) and \( Y \)) that can readily be proved by automatic provers:

```
hypothesis: \( Y' = \text{set}(Y, \text{all}, \text{get}(X, \text{all}) + 1) \)
goal: \( \text{get}(Y', \text{all}) = \text{get}(X, \text{all}) + 1 \)
```

In Section 2 we present a formalization of the non-aliasing rules enforced by the permission-based analysis, including proof of non-aliasing guarantees. In Section 3 we describe a concrete implementation of the analysis inside the open-source GNAT compiler for Ada which is part of GCC. We survey related works in Section 4 in particular with respect to Rust.
2 Alias Analysis

In Ada/SPARK code, the access to memory areas is given through \textit{paths} that start with an identifier (a variable name) and follow through record fields, array indices, or through a special field \texttt{all}, which corresponds to pointer dereferencing. In this paper, we only consider record and pointer types, and discuss the treatment of arrays in Section 3.

As an example, we use the following Ada type, describing singly linked lists where each node carries a boolean flag and a pointer to a shared integer value.

\begin{verbatim}
type List is record
  Flag : Boolean;
  Key : access Integer;
  Next : access List;
end record;
\end{verbatim}

Given a variable \texttt{A : List}, the paths \texttt{A.Flag}, \texttt{A.Key.all}, \texttt{A.Next.all.Key} are valid and their respective types are \texttt{Boolean}, \texttt{Integer}, and \texttt{access Integer} (a pointer to an \texttt{Integer}). The important difference between pointers and records in Ada is that—similarly to C—assignment of a record copies the values of fields, whereas assignment of a pointer only copies the address and creates an alias.

The alias analysis procedure runs after the type checking. The idea is to associate one of the four permissions—\texttt{RW}, \texttt{R}, \texttt{W} or \texttt{NO}—to each possible path (starting from the available variables) at each sequence point in the program.

The \texttt{read-only} permission \texttt{R} allows us to read any value accessible from the path: use it in a computation, or pass it as an \texttt{in} parameter in a procedure call. As a consequence, if a given path has the \texttt{R} permission, then each valid extension of this path also has it.

The \texttt{write-only} permission \texttt{W} allows us to modify memory occupied by the value: use it on the left-hand side in an assignment or pass it as an \texttt{out} parameter in a procedure call. For example, having a write permission for a path of type \texttt{List} allows us to modify the \texttt{Flag} field or to change the addresses stored in the pointer fields \texttt{Key} and \texttt{Next}. However, this does not necessarily give us the permission to modify memory accessible from those pointers. Indeed, to dereference a pointer, we must read the address stored in it, which requires the read permission. Thus, the \texttt{W} permission only propagates to path extensions that do not dereference pointers, i.e., do not contain additional \texttt{all} fields.

The \texttt{read-write} permission \texttt{RW} combines the properties of the \texttt{R} and \texttt{W} permissions and grants the full ownership of the path and every value accessible from it. In particular, the \texttt{RW} permission propagates to all valid path extensions including those that dereference pointers. The \texttt{RW} permission is required to pass a value as an \texttt{in-out} parameter in a procedure call.

Execution of program statements changes permissions. For example, allocating a new non-initialised memory area assigns the \texttt{W} permission to every value stored in this area. Thus, after the statement \texttt{P := new List}, the paths \texttt{P}, \texttt{P.all}, \texttt{P.all.Flag}, \texttt{P.all.Key}, and \texttt{P.all.Next} have permission \texttt{W}. All strict extensions of \texttt{P.all.Key} and \texttt{P.all.Next} receive permission \texttt{NO}, that is, no
procedure P1 (A, B: in out List) is begin
A := B;
B.Flag := True;
B.Key.all := 42;
end P1;

procedure P2 (A, B: in out Integer_Pointer) is begin
while B.all > 0 loop
A.all := A.all + 1;
B.all := B.all - 1;
A := B;
end loop;
end P2;

Fig. 1. Examples of potentially harmful aliasing.

permission. Indeed, since the pointers P.all.Key and P.all.Next are not initialised, neither reads nor writes under them make any sense. Another simple example of permission change is the procedure call: all out parameters must be assigned by the callee and receive the RW permission after the call.

The assignment statement is more complicated and several cases must be considered. If we assign a value that does not contain pointers (say, an integer or a pointer-free record), the whole value is copied into the left-hand side, and we only need to check that we have the appropriate permissions: W or RW for the left-hand side and R or RW for the right-hand side. However, whenever we copy a pointer, an alias is created. We want to make the left-hand side the new full owner of the value (i.e., give it the RW permission), and therefore, after the permission checks, we must revoke the permissions from the right-hand side, to avoid potentially harmful aliasing. The permission checks are also slightly different in this case, as we require the right-hand side to have the RW permission in order to move it to the left-hand side.

Let us now consider several simple programs and see how the permission checks allow us to detect potentially harmful aliasing.

Procedure P1 in Fig. 1 receives two in-out parameters A and B of type List. At the start of the procedure, all in-out parameters assume permission RW. In particular, this implies that each in-out parameter is separated from all other parameters (in fact, only the in parameters may alias each other). The first assignment copies the structure B into A. Thus, the paths A.Flag, A.Key, and A.Next are separated, respectively, from B.Flag, B.Key, and B.Next. However, the paths A.Key.all and B.Key.all are aliased, and A.Next.all and B.Next.all are aliased as well.

The first assignment does not change the permissions of A and its extensions: they retain the RW permission and keep the full ownership of their respective memory areas, even if the areas themselves have changed. The paths under B, however, must relinquish (some of) their permissions. The paths B.Key.all and B.Next.all as well as all their extensions get the NO permission, that is, lose both read and write permissions. This is necessary, as the ownership over their memory areas is transferred to the corresponding paths under A. The paths B, B.Key, and B.Next lose the read permission but keep the write-only W
permission. Indeed, we forbid reading from memory that can be altered through a concurrent path. However, it is allowed to “redirect” the pointers B.Key and B.Next, either by assigning those fields directly or by copying some different record into B. The field B.Flag is not aliased, nor has it aliased extensions, and thus retains the initial RW permission. This RW permission allows us to perform the assignment B.Flag := True on the next line.

The third assignment, however, is now illegal, since B.Key all does not have the write permission anymore. What is more, at the end of the procedure the in-out parameters A and B are not separated. This is forbidden, as the caller assumes that all out and in-out parameters are separated after the call just as they were before.

Procedure P2 in Fig. 1 receives two pointers A and B, and manipulates them inside a while loop. Since the permissions are assigned statically, we must ensure that at the end of a single iteration, we did not lose the permissions necessary for the next iteration. This requirement is violated in the example: after the last assignment A := B, the path B receives permission W and the path B.all, permission NO, as B.all is now an alias of A.all. The new permissions for B and B.all are thus weaker than the original ones (RW for both), and the procedure is rejected. Should it be accepted, we would have conflicting memory modifications from two aliased paths at the beginning of the next iteration.

2.1 µSPARK language

For the purposes of formal presentation, we introduce µSPARK, a small subset of SPARK featuring pointers, records, loops, and procedure calls. We present the syntax and semantics of µSPARK, and define the rules for static analysis of alias safety.

The data types of µSPARK are as follows:

\[
type ::= \text{Integer} \mid \text{Real} \mid \text{Boolean} \quad \text{scalar type} \\
\mid \text{access type} \quad \text{access type (pointer)} \\
\mid \text{ident} \quad \text{record type}
\]

Every µSPARK program starts with a list of record type declarations:

\[
record ::= type \text{ ident is record } field^* \text{ end} \\
field ::= ident : type
\]

We require all field names to be distinct. The field types must not refer to the record types declared later in the list. Recursive record types are allowed: a field of a record type R can contain pointers to R (written access R). We discuss the handling of array types in Section 3.
The syntax of $\mu$SPARK statements is defined by the following rules:

$$path ::= ident$$  
$$| path \cdot ident$$  
$$| path \cdot all$$

$$expr ::= path$$  
$$| 42 | 3.14 | True | False | ...$$  
$$| expr \ (+\ -\ |\ <\ |\ =\ |\ ...) expr$$  
$$| path^Access$$  
$$| null$$

$$stmt ::= path := expr$$  
$$| path := new type$$  
$$| if expr then stmt^* else stmt^* end$$  
$$| while expr loop stmt^* end$$  
$$| ident ( expr^* )$$

Following the record type declarations, a $\mu$SPARK program contains a set of mutually recursive procedure declarations:

$$procedure ::= procedure ident ( param^* ) is local^* begin stmt^* end$$

$$param ::= ident : ( in \mid in-out \mid out ) type$$

$$local ::= ident : type$$

We require all formal parameters and local variables in a procedure to have distinct names. A procedure call can only pass left-values (i.e., paths) for in-out and out parameters. The execution starts from a procedure named Main with the empty parameter list.

The type system for $\mu$SPARK is rather standard and we do not present it here in full. We assume that binary operators only operate on scalar types. The null pointer can have any pointer type access $\tau$. The dereference operator .all converts a pointer type access $\tau$ to $\tau$. The access operator 'Access applied to an l-value of type $\tau$ returns the corresponding pointer type access $\tau$. The allocation statement $p := new \tau$ requires the path $p$ to have type access $\tau$. In what follows, we only consider well-typed $\mu$SPARK programs.

On the semantic level, we need to distinguish the units of allocation, such as whole records, from the units of access, such as individual record fields. We use the term location to refer to the memory area occupied by an allocated value. We treat locations as elements of an abstract infinite set, and denote them with letter $\ell$. We use the term address to designate either a location, denoted $\ell$, or a specific component inside the location of a record, denoted $\ell.f.g$, where $f$ and $g$ are field names (assuming that at $\ell$ we have a record whose field $f$ is itself a record with a field $g$). A value is either a scalar, an address, a null pointer or a record, that is, a finite mapping from field names to values.

A $\mu$SPARK program is executed in the context defined by a binding $\Upsilon$ that maps variable names to addresses and a store $\Sigma$ that maps locations to values.
By a slight abuse of notation, we apply $\Sigma$ to arbitrary addresses, so that $\Sigma(\ell,f)$ is $\Sigma(\ell)(f)$, the value of the field $f$ of the record value stored in $\Sigma$ at $\ell$. Similarly, we write $\Sigma[\ell.f \mapsto v]$ to denote an update of a single field in a record, that is, $\Sigma[\ell \mapsto \Sigma(\ell)[f \mapsto v]]$.

We use big-step operational semantics and write $\Upsilon \cdot \Sigma \cdot s \Downarrow \Sigma'$ to denote that $\mu$SPARK statement $s$, when evaluated under binding $\Upsilon$ and store $\Sigma$, terminates with the state of the store $\Sigma'$. We extend this notation to sequences of statements $s$ in an obvious way. In this paper, we do not consider diverging statements.

The evaluation of expressions is effect-free and is denoted $\llbracket e \rrbracket_{\Upsilon \Sigma}$. We also need to evaluate l-values to the corresponding addresses in the store, written $\langle \langle p \rangle \rangle_{\Upsilon \Sigma}$, where $p$ is the evaluated path. Illicit operations, such as dereferencing a null pointer, cannot be evaluated and stall the execution (blocking semantics). In the formal rules below, $c$ stands for a scalar constant and $\odot$, for a binary operator:

$$
\langle x \rangle_{\Upsilon \Sigma} = T(x)
$$
$$
\langle p.f \rangle_{\Upsilon \Sigma} = \langle p \rangle_{\Upsilon \Sigma}.f
$$
$$
\langle p.{\text{all}} \rangle_{\Upsilon \Sigma} = \langle p \rangle_{\Upsilon \Sigma}
$$
$$
\langle p' \text{Access} \rangle_{\Upsilon \Sigma} = \langle p \rangle_{\Upsilon \Sigma}
$$
$$
\llbracket e_1 \odot e_2 \rrbracket_{\Upsilon \Sigma} = \llbracket e_1 \rrbracket_{\Upsilon \Sigma} \odot \llbracket e_2 \rrbracket_{\Upsilon \Sigma}
$$
$$
\llbracket \text{null} \rrbracket_{\Upsilon \Sigma} = \text{null}
$$

Allocation adds a fresh address to the store, mapping it to a default value for the corresponding type: 0 for Integer, False for Boolean, null for the access types, and for the record types, a record value where each field has the default value. Notice that since pointers are initialised to null, there is no deep allocation. We write $\square_\tau$ to denote the default value of type $\tau$.

The evaluation rules are given in Figure 2. In the (E-call) rule, we evaluate the procedure body in the dedicated context $T_p \cdot \Sigma_p$. This context binds the in parameters to fresh locations containing the values of the respective expression arguments, binds the in-out and out parameters to the addresses of the respective l-value arguments, and allocates memory for the local variables. For simplicity, we do not reclaim memory on return from a procedure call, and thus avoid dangling pointers. In Ada and SPARK, this issue is handled separately, using scope-based memory pools, and does not need to be addressed by our analysis procedure.

### 2.2 Access policies, transformers, and alias safety rules

We denote paths with letters $p$ and $q$. We write $p \sqsubseteq q$ to denote that $p$ is a strict prefix of $q$ or, equivalently, $q$ is a strict extension of $p$. In what follows, we always mean strict prefixes and extensions, unless explicitly said otherwise.

In the typing context of a given procedure, a well-typed path is said to be deep if it has an extension of an access type, otherwise it is called shallow. We extend these notions to types: a type $\tau$ is deep (resp. shallow) if and only if a $\tau$-typed path is deep (resp. shallow). In other words, a path or a type is deep if a pointer can be reached from it, and shallow otherwise. For example, the List type in Section 2 is a deep type, and so is access Integer, whereas any scalar type or any record with scalar fields only is shallow.
A dereferences as since they all create an additional pointer dereference by passing through compute an statement. For each sequence point in a given µ one of the four permissions occur as sub-expressions: in an expression X.f.g + Y.h dating the policy for every path in the expression. This only includes paths that RW Π We write T from application of transformer abbreviation for Π and extensions. Symbolically, we write Σp, q = Σ[ℓa1 → va1,..., ld1 → rd1,...]. Σp = Σ[ℓa1 → va1,..., ld1 → rd1,...] Tp · Σp · s ↓ Σ′ (E-CALL)

procedure P (a1 : in τa1; ...; b1 : in-out τb1; ...; c1 : out τc1; ...)

is d1 : τd1; ... begin s end is declared in the program

ℓa1,...,ld1,... ⊆ dom Σ [ea1]Σ = va1,...,

Σp = [a1 → ℓa1,..., b1 → (p_a1)Σ, ..., c1 → (q_c1)Σ, ..., d1 → ℓd1,...] Tp · Σp · s ↓ Σ′ (E-CALL)

Fig. 2. Semantics of µSPARK (terminating statements).

An extension q of a path p is called a near extension if it has as many pointer dereferences as p, otherwise it is a far extension. For instance, given a variable A of type List, the paths A.Flag, A.Key, and A.Next are the near extensions of A, whereas A.Key.all, A.Next.all, and their extensions are far extensions, since they all create an additional pointer dereference by passing through all.

We say that sequence points are the program points before or after a given statement. For each sequence point in a given µSPARK program, we statically compute an access policy: a partial function that maps each well-typed path to one of the four permissions: RW, R, W, and NO, which form a diamond lattice: RW > R|W > NO. We denote permissions with π and access policies with Π.

Permission transformers modify policies at a given path, as well as its prefixes and extensions. Symbolically, we write Π →p Π′ to denote that policy Π′ results from application of transformer T to Π at path p. We write Π →p Π′ as an abbreviation for Π →p Π′ (that is, for some Π″, Π →p Π″ →p Π′). We write Π →p,q Π′ as an abbreviation for Π →p,q Π′.

Permission transformers can also apply to expressions, which consists in updating the policy for every path in the expression. This only includes paths that occur as sub-expressions: in an expression X.f.g + Y.h, only the paths X.f.g
and only verifies that a given path \( p \) possesses the
check that the right-hand side is initially the sole owner of the copied value
over the copied value to the left-hand side. If we copy a value of a shallow
¯
and ‘borrow’, which we define and explain below.

Let us start with the (P-Assign) rule. Assignments grant the full ownership
over the copied value to the left-hand side. If we copy a value of a shallow
type, we merely have to ensure that the right-hand side has the read permission.
Whenever we copy a deep-typed value, aliases may be created, and we must
check that the right-hand side is initially the sole owner of the copied value
(that is, possesses the RW permission) and revoke the ownership from it.

To define the ‘move’ transformer that handles permissions for the right-hand
side of an assignment, we need to introduce several simpler transformers.

**Definition 1.** Permission transformer check \( \pi \) does not modify the access policy and only verifies that a given path \( p \) has permission \( \pi \) or greater. In other words, \( \Pi \xrightarrow{\text{check } \pi} \Pi' \) if and only if \( \Pi(p) \geq \pi \) and \( \Pi = \Pi' \). This transformer also applies to expressions: \( \Pi \xrightarrow{\text{check } \pi} \Pi' \) states that \( \Pi \xrightarrow{\text{check } \pi} \Pi' (= \Pi) \) for every path \( p \) occurring in \( \epsilon \).

---

Fig. 3. Alias safety rules for statements.
Definition 2. Permission transformer fresh π assigns permission π to a given path p and all its extensions.

Definition 3. Permission transformer cut assigns restricted permissions to a deep path p and its extensions: the path p and its near deep extensions receive permission W, the near shallow extensions keep their current permissions, and the far extensions receive permission NO.

Going back to the procedure P1 in Fig. 1, the change of permissions on the right-hand side after the assignment A := B corresponds to the definition of ‘cut’. In the case where the right-hand side of an assignment is not simply a variable, but a deep path or a 'Access expression, we also need to change permissions of the prefixes, to reflect the ownership transfer.

Definition 4. Permission transformer block propagates the loss of the read permission from a given path to all its prefixes. Formally, it is defined by the following rules, where x stands for a variable and f for a field name:

\[
\begin{align*}
\Pi & \xrightarrow{\text{block}}_x \Pi \\
\Pi(p) & = \text{NO} \\
\Pi & \xrightarrow{\text{block}}_{p.f} \Pi
\end{align*}
\]

Definition 5. Permission transformer drop propagates the loss of both read and write permissions from a given path to its prefixes up to the first pointer, and propagates the loss of the read permission afterwards:

\[
\begin{align*}
\Pi & \xrightarrow{\text{drop}}_x \Pi \\
\Pi[p \mapsto W] & \xrightarrow{\text{block}}_{p.all} \Pi' \\
\Pi[p \mapsto \text{NO}] & \xrightarrow{\text{drop}}_{p.f} \Pi'
\end{align*}
\]

Definition 6. Permission transformer move applies to expressions:

- if e has a shallow type, then \(\Pi \xrightarrow{\text{move}}_e \Pi' \iff \Pi \xrightarrow{\text{check RW} \land \text{cut} \land \text{block}} \Pi'\);
- if e is a deep path p, then \(\Pi \xrightarrow{\text{move}}_{p} \Pi' \iff \Pi \xrightarrow{\text{check RW} \land \text{cut} \land \text{block}} \Pi'\);
- if e is a 'Access expression, then \(\Pi \xrightarrow{\text{move}}_{p} \Pi' \iff \Pi \xrightarrow{\text{check RW} \land \text{fresh} \land \text{drop} \land \text{block}} \Pi'\);
- if e is null, then \(\Pi \xrightarrow{\text{move}}_{\text{null}} \Pi' \iff \Pi' = \Pi\).

To further illustrate the ‘move’ transformer, let us consider two variables P and Q of type access List and an assignment P := Q.all.Next. We assume that Q and all its extensions have full ownership (RW) before the assignment. We apply the second case in the definition of ‘move’ to the deep path Q.all.Next. The ‘check RW’ condition is verified, and the ‘cut’ transformer sets the permission for Q.all.Next to W and the permission for Q.all.Next.all and all its extensions to NO. Indeed, P.all becomes an alias of Q.all.Next.all and steals the full ownership for this memory area. However, we still can reassign
Q.all.Next to a different address. Moreover, we still can write some new values into Q.all or Q, without compromising safety. This is enforced by the application of the ‘block’ transformer at the end. We cannot keep the read permission for Q or Q.all, since it implies the read access to the data under Q.all.Next.all.

Now, let a variable R have type access Boolean and consider the assignment R := Q.all.Flag’Access. We apply the third case in the definition of ‘move’. Assuming once again that Q has full ownership over its value, the ‘check RW’ condition for Q.all.Flag is verified. Since the ownership of this Boolean value is now transferred to R.all, we must revoke all permissions from Q.all.Flag (and its extensions, if it had any), which is enforced by ‘fresh NO’. Moreover, since writing into the record Q.all overwrites the Flag field, we must also revoke all permissions from Q.all. This is done by the ‘drop’ transformer. Notice that the permissions for Q.all.Key and Q.all.Next are not affected: we can still read and modify those fields, as they are not aliased with other paths. Furthermore, modifying the pointer Q itself is allowed, which is why ‘drop’ becomes ‘block’ after rising past all.

Finally, we need to describe the change of permissions on the left-hand side of an assignment, in order to reflect the gain of the full ownership. The idea is that as soon as we have the full ownership for each field of a record, we can assume the full ownership of the whole record, and similarly for pointers.

**Definition 7.** Permission transformer lift propagates the RW permission from a given path to its prefixes, wherever possible:

\[
\begin{align*}
\Pi & \xrightarrow{\text{lift}_p} \Pi' \\
\forall q \sqsubseteq p. \Pi(q) = RW & \quad \Pi[p \mapsto RW] \xrightarrow{\text{lift}_p} \Pi'
\end{align*}
\]

In the (P-assign) rule, we revoke the permissions from the right-hand side of an assignment before granting the ownership to the left-hand side. This is done in order to prevent creation of circular data structures. Consider an assignment A.Next := A’Access, where A has type List. According to the definition of ‘move’, path A and all its extensions receive permission NO. This makes the left-hand side A.Next fail the write permission check.

Allocations p := new τ are handled by the (P-alloc) rule. As long as the memory area under the pointer p is not explicitly initialised by the program code, no read permission is granted for p, nor its extensions and prefixes. Moreover, since the pointer fields (if any) in the allocated memory are not accessible yet, no permission at all can be given for the far extensions of p. This is enforced by the ‘cut’ transformer.

In a conditional statement, the policies at the end of the two branches are merged selecting the most restrictive permission for each path. Loops require that no permissions are lost at the end of a loop iteration, compared to the entry, as explained above for procedure P2 in Fig. 1.
Procedure calls guarantee to the callee that every argument with mode in, in-out, or out has at least permission R, RW or W, respectively. To ensure the absence of potentially harmful aliasing, we revoke the necessary permissions using the ‘observe’ and ‘borrow’ transformers.

**Definition 8.** Permission transformer `borrow` assigns permission NO to a given path `p` and all its prefixes and extensions.

**Definition 9.** Permission transformer `freeze` removes the write permission from a given path `p` and all its prefixes and extensions. In other words, freeze assigns to each path `q` comparable to `p` the minimum permission `\Pi(q) \land R`.

**Definition 10.** Permission transformer `observe` applies to expressions:
- if `e` has a shallow type, then `\Pi_{\text{observe}} \rightarrow e \Pi' \Leftrightarrow \Pi' = \Pi`;
- if `e` is a deep path `p`, then `\Pi_{\text{observe}} \rightarrow e \Pi' \Leftrightarrow \Pi_{\text{freeze}} \rightarrow_p \Pi'`;
- if `e` is `Access`, then `\Pi_{\text{observe}} \rightarrow e \Pi' \Leftrightarrow \Pi_{\text{freeze}} \rightarrow_p \Pi'`;
- if `e` is `null`, then `\Pi_{\text{observe}} \rightarrow e \Pi' \Leftrightarrow \Pi' = \Pi`.

We remove the write permission from the deep-typed in parameters using the ‘observe’ transformer, in order to allow aliasing between the read-only paths. As for the in-out and out parameters, we transfer the full ownership over them to the callee, which is reflected by dropping every permission on the caller’s side using ‘borrow’.

In the (P-CALL) rule, we revoke permissions right after checking them for each parameter. In this way, we cannot pass, for example, the same path as an in and in-out parameter in the same call. Indeed, the ‘observe’ transformer will remove the write permission, which is required by ‘check RW’ later in the transformer chain. At the end of the call, the callee transfers to the caller the full ownership over each in-out and out parameter.

We apply our alias safety analysis to each procedure declaration. We start with an empty access policy, denoted `\emptyset`. Then we fill the policy with the permissions for the formal parameters and the local variables and check the procedure body. At the end, we verify that every in-out and out parameter has the RW permission. Formally, this is expressed with the following rule:

\[
\begin{align*}
\emptyset & \xrightarrow{\text{fresh \ R}} \gamma_{a_1, \ldots, \gamma} \xrightarrow{\text{fresh \ RW}} \gamma_{b_1, \ldots, \gamma} \xrightarrow{\text{fresh \ W \ cut}} \gamma_{c_1, \ldots, \gamma} \xrightarrow{\text{fresh \ W \ cut}} \gamma_{d_1, \ldots, \gamma} \\
\Pi' \cdot \delta & \rightarrow \Pi'' \\
\Pi''(b_1) & = \cdots = \Pi''(c_1) = \cdots = \text{RW}
\end{align*}
\]

\text{procedure } P \langle a_1 : \tau_{a_1}; \ldots; b_1 : \text{in-out } \tau_{b_1}; \ldots; c_1 : \text{out } \tau_{c_1}; \ldots \rangle \\
\text{is } d_1 : \tau_{d_1}; \ldots \text{ begin } s \text{ end } \text{ is alias-safe}

We say that a \(\mu\)SPARK program is alias-safe if all its procedures are.
2.3 Soundness

As the end of the analysis, an alias-safe program has an access policy associated to each sequence point in it. We say that an access policy \(\Pi\) is consistent whenever it satisfies the following conditions for all valid paths \(\pi, \pi.f, \pi.\text{all}\):

\[
\begin{align*}
\Pi(\pi) = \text{RW} \Rightarrow \Pi(\pi.f) = \text{RW} & \quad \Pi(\pi) = \text{RW} \Rightarrow \Pi(\pi.\text{all}) = \text{RW} \quad (1) \\
\Pi(\pi) = \text{R} \Rightarrow \Pi(\pi.f) = \text{R} & \quad \Pi(\pi) = \text{R} \Rightarrow \Pi(\pi.\text{all}) = \text{R} \quad (2) \\
\Pi(\pi) = \text{W} \Rightarrow \Pi(\pi.f) \geq \text{W} & \quad \Pi(\pi) = \text{W} = \Rightarrow \Pi(\pi.\text{all}) = \text{W} \quad (3)
\end{align*}
\]

These invariants correspond to the informal explanations given in Section 2. Invariant (1) states that the full ownership over a value propagates to all values reachable from it. Invariant (2) states that the read-only permission must also propagate to all extensions. Indeed, a modification of a reachable component can be observed from any prefix. Invariant (3) states that write permission over a record value implies a write permission over each of its fields. However, the write permission does not necessarily propagate across pointer dereference.

Lemma 1 (Policy consistency). The alias safety rules in Fig. 3 preserve policy consistency.

When, during an execution, we arrive at a given sequence point with the set of variable bindings \(\Upsilon\), store \(\Sigma\), and statically computed and consistent access policy \(\Pi\), we say that the state of the execution respects the Concurrent Read, Exclusive Write condition (CREW), if and only if for any two distinct valid paths \(p\) and \(q\), \(\langle p \rangle^\Upsilon_\Sigma = \langle q \rangle^\Upsilon_\Sigma \land \Pi(\pi) \geq \text{W} \Rightarrow \Pi(\pi) = \text{NO} \).

The main result about the soundness of our approach is as follows:

Theorem 1 (Soundness). A terminating evaluation of a well-typed alias-safe \(\mu\)SPARK program respects the CREW condition at every sequence point.

The full proof, for a slightly different definition of \(\mu\)SPARK, is given in [3].

The argument proceeds by induction on the evaluation derivation, following the rules in Figure 2. The only difficult cases are assignment, where the required permission withdrawal is ensured by the ‘move’ transformer, and procedure call, where the chain of ‘observe’ and ‘borrow’ transformers, together with the corresponding checks, on the caller’s side, ensures that the CREW condition is respected at the beginning of the callee.

3 Implementation

We have implemented the permission rules in the GNAT compiler for Ada, which is part of GCC. The implementation consists in 3700 lines in Ada. The analysis is triggered by using the debug switch -gnatdF and setting the context for formal verification of SPARK code with switch -gnatd.F and aspect SPARK_Mode.

Access policies are infinitely big in presence of recursive types, which we took into account in our implementation with a lazy implementation of permission trees. Permission trees start with a depth of one, and are expanded on demand.
The permission rules presented in Section 2 only address a subset of SPARK. Complete SPARK differs from \( \mu \)SPARK on several points, which have been taken into account when designing the full set of rules. For arrays, permission rules are adapted to apply to all elements, without taking into account the exact index of that element, which may not be known statically in general. Besides procedures, SPARK has functions, which return values and cannot perform side-effects. Functions take only parameters of mode \texttt{in} and can be called inside expressions. Our permission rules are augmented with an implicit move of the values returned by functions, which allows us to support constructors.

In our formalization, we considered that every shallow \texttt{in} parameter is passed by-copy, which is not the case in SPARK. In our implementation, we correctly distinguish parameters of by-copy types (typically scalars) and parameters which may be passed by-copy or by-reference, which we treat as deep parameters (except for function \texttt{in} parameters, as functions cannot have side effects, and in particular cannot write in their parameters even under dereference).

Loops in SPARK also differ on two accounts from the loops in our formalization: besides ‘while’ loops, SPARK also defines ‘for’ loops and plain loops (with no exit condition), and exit statements inside a loop allow exiting any enclosing loop. In our implementation, we take these into account to create the correct merged context for analyzing the code after the loop.

The analysis has been tested on four test suites, including the regression testsuite of the GNAT compiler (17041 tests [4]) and the regression testsuite of the SPARK product (2087 tests), which detected 30 regressions, most of them related to object oriented features that the current analysis is not able to handle. We also wrote a dedicated testsuite [5] containing 20 tests inspired by the examples given in the Rust borrow-checker documentation [6].

4 Related Works

4.1 Comparison with Rust

The following section compares Rust and SPARK on some features and constructs that seem relevant to the authors. Both prevent harmful aliasing in the source code, whereas different design choices affect their respective expressiveness. In SPARK, an additional limitation was the choice not to add any annotation or keyword to the language.

The main differences come from the fact that SPARK uses other compiler passes to handle many safety features, whereas they are handled directly by Rust’s borrow-checker and type-checker. A benefit of our work is that it unambiguously defines the rules for SPARK while there exists no official document specifying Rust’s borrow-checker, in all details, especially as Rust continues to go through significant evolutions, such as non-lexical lifetimes [7]. We must note, however, a significant recent effort to provide a rigorous formal description of the foundations of Rust [8].

In SPARK, the duration of borrows is limited to the duration of procedure calls. The mechanism of \textit{renaming} in SPARK can be used to shorten long paths,
and thus it is less important to be able to create several local copies of the same
deep variable. It turns out, however, that this restriction forbids traversing a
linked data structure with only R permission, and that even with RW permission
the structure needs to be reconstructed after traversal. We are working on the
rules for local application of borrow and observe inside a block to allow these.

Rust has sophisticated lifetime checks, allowing to precisely control the du-
rational of a borrowed pointer. In SPARK, similar checks are implemented as a
separate analysis pass of GNAT compiler, with their own set of rules that are
less expressive: the lifetime of a pointer is limited to the block in which its type
has been declared.

There are no null pointers in Rust, whereas they are allowed in SPARK.
Dereferencing a null pointer is a runtime error, and the null-pointer safety must
be proved by a separate analysis by the SPARK verification tools [9].

Another difference is that mutable borrows in Rust guarantee the ownership
of the underlying memory at any time, whereas in SPARK borrow guarantees
ownership only at the entry and exit of the procedure. In particular, it is not
possible in Rust to move out a borrowed variable without assigning it in the
same statement. In SPARK, any borrowed variable can be moved as long as it
is assigned before the end of the procedure, which grants it the RW permission.
This allows us to directly implement the swap procedure in SPARK, whereas
the Rust implementation relies on unsafe operations:

\[
\begin{align*}
\text{procedure } & \text{Swap (X, Y : in out T) is} \\
& \text{Temp : T := Y; } \quad \text{-- Move Y. X:RW, Y:W, Temp:RW} \\
& \text{begin} \\
& \quad Y := X; \quad \quad \text{-- Move X. X:W, Y:RW, Temp:RW} \\
& \quad X := \text{Temp; } \quad \quad \text{-- Move Temp. X:RW, Y:RW, Temp:W} \\
& \text{end Swap;} \quad \quad \text{-- Both arguments X and Y are RW.}
\end{align*}
\]

Finally, it is impossible to implement cycling constructs in SPARK (the right-
hand side of an assignment cannot be an ancestor of the left-hand side). Rust
has a similar limitation and requires some workarounds like reference counting to
implement structures like graphs [10]. Nevertheless, both Rust and SPARK allow
compiling with parts of code written in an unsafe superset of the language. This
is the case for some standard library containers (hash tables, iterators, smart
pointers, etc.) that have an interface specified in SPARK and Rust, and can be
called safely from safe Rust/SPARK.

4.2 Other Related Works

Permission-based programming languages generalize the issue of avoiding harm-
ful aliasing to the more general problem of preventing harmful sharing of re-
sources (memory, but also network connections, files, etc.). Cyclone and Rust
achieve absence of harmful aliasing by enforcing an ownership type system on
the memory pointed to by objects [11][12].

Dafny associates each object with its \textit{dynamic frame}, the set of pointers
that it owns [13]. This dynamic version of ownership is enforced by modeling
the ownership of pointers in logic, generating verification conditions to detect violations of the single-owner model, and proving them using SMT provers. In Spec#, ownership is similarly enforced by proof, to detect violations of the so-called Boogie methodology [14].

Separation logic [15] is an extension of Hoare-Floyd logic that allows reasoning about pointers. In general, it is difficult to integrate into automated deductive verification: in particular, it is not directly supported by SMT provers.

In our work, we use a permission-based mechanism for detecting potentially harmful aliasing, in order to make the presence of pointers transparent for automated provers. In addition, our approach does not require additional user annotations, that are required in some of the previously mentioned techniques. We thus expect to achieve high automation and usability, which was our goal for supporting pointers in SPARK.

5 Conclusion

In this paper, we have presented anti-aliasing rules that allow supporting pointers in SPARK. We showed a systematic analysis that allows a wide range of use cases with pointers and dynamic allocation. To the best of our knowledge, this is a novel approach for controlling aliasing introduced by arbitrary pointers in a programming language supported by proof. Our approach does not require user annotations or proof of verification conditions, which makes it much simpler to adopt. Moreover, we provided a formalization of our rules on a subset of SPARK in order to mathematically prove the safety of our analysis. Finally, we compared our method to the Rust language which provides a similar analysis.

More work needs to be done to fully support pointers in SPARK. Both flow analysis and proof need to be adapted to account for the presence of pointers. This work has started and is expected to be completed by the end of 2018. This will make it possible to use formal verification with SPARK on industrial programs with pointers, something that was long believed to be impossible.

We also need to extend our formalism and proof to non-terminating executions. For that purpose, we can provide a co-inductive definition of the big-step semantics and perform a similar co-inductive soundness proof, as described by Leroy and Grall [16].

The formal study of the constructs that can be implemented using this borrow-checker has not been discussed in this paper. Such a study could allow a formalization of anti-aliasing analyses through classes of expressiveness. Another long-term goal would be extending our analysis so that it could handle automatic reclamation, parallelism, initialization and lifetime checks, instead of relying on external checks.

The proposed anti-aliasing rules are being discussed by the Ada Rapporteur Group for inclusion in the next version of Ada [17].
References

1. McCormick, J.W., Chapin, P.C.: Building High-Integrity Applications with SPARK. Cambridge University Press (2015)
2. Kirchner, F., Kosmatov, N., Prevosto, V., Signoles, J., Yakobowskib, B.: Framac: A software analysis perspective. Formal Aspects of Computing 27(3) (May 2015) 573–609
3. Jaloyan, G.A.: Safe pointers in SPARK 2014 (2017) https://arxiv.org/abs/1710.07047
4. ACATS: Test suite http://www.ada-auth.org/acats-files/4.1/ACATS41.ZIP
5. Georges-Axel Jaloyan: SPARK aliasing test suite. https://github.com/GAJaloyan/SPARKtestsuite
6. Rust community: Documentation for Rust’s borrow-checker. https://github.com/rust-lang/rust/blob/master/src/librustc_borrowck/borrowck/README.md
7. Rust community: Non-lexical lifetimes. https://github.com/rust-lang/rfcps/pull/2094
8. Ralf Jung, Jacques-Henri Jourdan, Robbert Krebbers, Derek Dreyer: RustBelt: Securing the foundations of the Rust programming language. 2(Issue POPL, 66) (2018)
9. AdaCore, Thales: Implementation Guidance for the Adoption of SPARK, http://www.adacore.com/uploads/technical-papers/ePDF-ImplementationGuidanceSPARK.pdf
10. Rust: Graphs and arena allocation https://github.com/nrc/r4cppp/blob/master/graphs/README.md
11. Grossman, D., Morrisett, G., Jim, T., Hicks, M., Wang, Y., Cheney, J.: Region-based memory management in Cyclone. SIGPLAN Not. 37(5) (May 2002) 282–293
12. Balasubramanian, A., Baranowski, M.S., Burtsev, A., Panda, A., Rakamaric, Z., Ryzhyk, L.: System programming in Rust: Beyond safety. SIGOPS Oper. Syst. Rev. 51(1) (September 2017) 94–99
13. Leino, K.R.M. In: Dafny: An Automatic Program Verifier for Functional Correctness. Springer Berlin Heidelberg, Berlin, Heidelberg (2010) 348–370
14. Barnett, M., Chang, B.Y.E., DeLine, R., Jacobs, B., Leino, K.R.M. In: Boogie: A Modular Reusable Verifier for Object-Oriented Programs. Springer Berlin Heidelberg, Berlin, Heidelberg (2006) 364–387
15. Reynolds, J.C.: Separation logic: A logic for shared mutable data structures. In: Proceedings of the 17th Annual IEEE Symposium on Logic in Computer Science. LICS ’02, Washington, DC, USA, IEEE Computer Society (2002) 55–74
16. Leroy, X., Grall, H.: Coinductive big-step operational semantics. Information and Computation 207(2) (Feb 2009) 284–304
17. AdaCore: Access value ownership and parameter aliasing (2018) http://www.ada-auth.org/cgi-bin/cvsweb.cgi/ai12s/ai12-0240-1.txt