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DATA TYPES AS ALGORITHMS (*)

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Abstract. — This paper presents a simplified construction of an algorithm space. The intuitive notion of an algorithm collection goes back to Nolin 1974. Our goal is here to formalize and simplify this presentation of algorithm theory. The introduction of a computability notion was induced by the utilization of bundle theory. We build an enumerable algorithm space where every object is computable.


INTRODUCTION

This paper presents a simplified construction of an algorithm space [6, 3]. The idea and the terminology (algorithm) go back to Nolin 1974 [6], who wanted a semantics for a programming language with type declarations. (It must be pointed out here that Nolin's "algorithms" are not algorithms in the usual sense, i.e. a step-by-step description of a computation. They are in fact much closer to generalized types, or algebraic sorts.)

However, a proof of the conceptual reasonableness of the idea of an algorithm space, or in other words, a mathematical proof of the existence of such a space, was missing. Such a proof was first given in a restricted setting it [2]. It used bundle theory, which, as shown in [3], underlies a significant part of programming language semantics model theory.

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In [8], which is a shorter version of [7], Nolin and Le Berre present a simplified version of this proof, where they use “solely elementary set theory properties” [7]. Technically, this amounts to abandon the upper bundle structure used in [3]. Unfortunately, we discovered [4, 5] an error in a crucial step of their argument: the representation of every normal self-function \( f : E \rightarrow E \) as an element of the domain \( E \), which makes the entire proof unsound. The falsehood of their theorem follows from a careful analysis of the representation used for threshold functions:

\[
f_{xy} : \quad z \rightarrow \text{if } z \leq x \text{ then } y \text{ else } T,
\]

which turns out to be identical to the representation used in [2, 3].

In this paper a simplified construction of an algorithm space is presented. The intent of this construction is to enrich and simplify Nolin's original presentation of algorithm theory.

The simplification we propose is a better mastering of algorithm “collections” cardinality: in fact we describe an enumerable algorithm collection. The enrichment is the introduction of computability notions: all our algorithms will be computable in some sense. And this makes the whole algorithm space itself effectively presentable.

1. **BUNDLE STRUCTURES OVER \( \mathcal{P}(\mathbb{N}) \)**

Let \( \mathbb{N} \) be the set of integers and \( E = \mathcal{P}(\mathbb{N}) \) be the set of subsets of \( \mathbb{N} \). We have:

\[
\forall x \in E, \quad x = \bigcup_{m \in x} \{m\} = \bigcup_{a_n \subseteq x} a_n;
\]

if we define an enumeration \( a : \mathbb{N} \rightarrow E \):

\[
a : \quad n \rightarrow \{n-1\} \quad \text{if } n \neq 0; \quad a_0 = \emptyset \quad \text{where } \emptyset \text{ is the empty set}.
\]

This gives an elementary monic ordered bundle structure over \( E \) [3].

The spectra are defined by:

\[
s : \quad E \rightarrow \mathcal{P}(E); \quad x \mapsto \{a_n : a_n \subseteq x\}.
\]

The elements \( a_n \) belong to their own spectrum and are therefore called rationals. They constitute the kernel of the bundle, which means that every
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\( x \in E \) can be obtained as the union of some well-chosen rationals [in fact the elements of the spectrum of \( x \), \( s(x) \)]. The function \( a : \mathbb{N} \to E, n \to a_n \) gives an enumeration of the kernel. The limit function is simply the set-theoretical union of elements of \( E \). This bundle structure will be called the lower bundle structure of \( E \).

On the other hand we also have:

\[ \forall x \in E, \quad x = \bigcap_{x \in b_m} b_m \]

if we define an enumeration:

\[ b : \quad \mathbb{N} \to E, \quad m \to b_m = \bigcap \{ k_0, k_1, \ldots, k_{p-1} \}, \]

where:

\[ m = \sum_{i < p} 2^k_i, \quad k_0 < k_1 < \ldots < k_{p-1}. \]

(Thus we use the dyadic expansion of integer \( m \).) Notice that \( b_0 = \mathbb{N} \).

The set of \( b_m \)'s is closed under finite intersection, and each \( b_m \) verifies the following algebraicity property:

For any descending chain \( \{ x_i \}_{i \in I} \) of elements of \( E \):

\[ b_m \supseteq \bigcap_{i} x_i \quad \Rightarrow \quad \exists i \ b_m \supseteq x_i. \]

The spectrum function:

\[ s : \quad E \to \mathcal{P}(E); \]

\[ x \to s(x) = \{ b_m : b_m \supseteq x \}, \]

defines an algebraic monic ordered bundle for the inverse inclusion over \( E \). Every \( b_m \) is rational, and the kernel of the bundle is exactly the set of all \( b_m \)'s.

The spectra are obviously closed under finite intersection, and the limit function is simply the set-theoretical intersection of elements of \( E \).

This structure will be called the upper bundle structure of \( E \).

Indeed the set of \( b_m \)'s is exactly the set of all co-finite subsets of \( \mathbb{N} \). Every co-finite set is recursive, thus recursively enumerable. Furthermore the relations \( a_n \subseteq a_m, b_n \subseteq b_m, a_n \subseteq b_m, b_m = b_n \cap b_p \) are all recursive in the indices \( m, n, p \).
2. COMPUTABILITY IN \( E = \mathcal{P}(\mathbb{N}) \)

2.1. Computable elements of \( \mathcal{P}(\mathbb{N}) \)

We say that an element \( x \in E \) is \textit{inf-computable}, i.e., computable when considered inside the lower bundle structure, if and only if the set \( \{ n : a_n \subseteq x \} \) is recursively enumerable in the index \( n \). We see at once that:

\[ x \in E \text{ is inf-computable} \iff x \text{ is r.e.} \]

We also define: \( x \in E \) is \textit{sup-computable} if and only if the set \( \{ m : x \subseteq b_m \} \) is recursively enumerable in the index \( m \).

**Fact 1:** \( x \in E \) is sup-computable \( \iff \) x is computable \( \iff \) x is inf-computable \( \iff \) x is r.e.

**Proof:**

\[ \{ m : x \subseteq b_m \} = \{ m : x \subseteq \bigcap \{ k_0, k_1, \ldots, k_{p-1} \}, m = \sum_{i \leq p} 2^{k_i}, k_i \text{ all different} \}, \]

\[ x \subseteq \bigcap \{ k_0, k_1, \ldots, k_{p-1} \} \iff \{ k_0, k_1, \ldots, k_{p-1} \} \subseteq x; \]

i.e., we have to enumerate all finite subsets of \( \bigcap \ x \). In fact one can show that \( \forall y \subseteq \mathbb{N} \) if \( \mathcal{P}_{\text{fin}}(y) \) is the set of finite subsets of \( y \), then:

\[ y \text{ is r.e.} \iff \mathcal{P}_{\text{fin}}(y) \text{ is r.e.} \]

For:

\[ y \text{ r.e.} \Rightarrow \mathcal{P}_{\text{fin}}(y) \text{ r.e.: we take a recursive enumeration of } y, \text{ and we enumerate all finite subsets of } y \text{ we can construct.} \]

\[ \mathcal{P}_{\text{fin}}(y) \text{ r.e.} \Rightarrow y \text{ r.e.: we take a recursive enumeration of } \mathcal{P}_{\text{fin}}(y) \text{ and we evaluate, in an effective manner, the cardinal of each (finite) subset of } y. \text{ If this cardinal is one, the subset is added to the enumeration of } y, \text{ otherwise we discard and consider the next one.} \]

Whence the three equivalences. \( \Box \)

**Definition:** \( x \in E \) is \textit{computable} if and only if \( x \) is inf-computable and \( x \) is sup-computable.

In a programming language, the availability of a ground data type, say \textit{integer}, amounts to the availability of a procedure with one variable, \( \text{int}(x) \), such that, for any input data \( a \), the call \( \text{int}(a) \) returns the value \textit{true} if \( a \in \mathbb{N} \) and \textit{false} otherwise. This is exactly realized by the recursive subset of \( \mathbb{N} \). More explicitly:
LEMMA 2: \( \forall x \in E = \mathcal{P}(\mathbb{N}) \):

(i) \( x \) is inf-computable \( \iff \) \( x \) is r. e.

(ii) \( x \) is sup-computable \( \iff \mathcal{C} \ x \) is r. e.

(iii) \( x \) is computable \( \iff x \) is recursive. \( \square \)

As a summary:

1. The computable elements of \( E \) are exactly the recursive subsets of \( \mathbb{N} \).

2. \( E \) is an elementary monic ordered bundle for \( \subseteq \), with \( \emptyset \) and the singletons as rational elements and the r.e. sets included in \( \mathbb{N} \) as inf-computable elements. The set of inf-computable (resp. computable) elements forms a sub-bundle of \( E \).

3. \( E \) is an algebraic bundle for \( \supseteq \), with the cofinite subsets as rational elements, and has as sup-computable elements (i.e., computable for this bundle) all subsets \( \subseteq \mathbb{N} \) whose complementary is r.e. The set of sup-computable (resp. computable) elements of \( E \) forms an algebraic sub-bundle of \( E \).

2.2. Computable sequences

We define computations in \( E \). We call them computable sequences.

DEFINITION: A sequence \( \{x_p\}_{p \in \mathbb{N}} \) of elements of \( E \) is said to be sup-computable (resp. inf-computable) if and only if there exists a recursive function \( \psi : \mathbb{N}^2 \rightarrow \mathbb{N} \), such that for any \( p \in \mathbb{N} \), \( \psi(p, .) \) is an enumeration of the (indices of the) spectrum of \( x_p \) for the upper bundle (resp. the lower bundle) structure of \( E \). \( \square \)

As a consequence, every term \( x_p \) of a computable sequence \( \{x_p\}_{p \in \mathbb{N}} \) is computable according to the bundle structure considered (i.e., \( x_p \) is inf-computable if the sequence is inf-computable; and \( x_p \) is sup-computable if the sequence is sup-computable).

There is an important fact about the compatibility of the computability and bundle notions in \( E \).

LEMMA 3: For any sequence \( \{x_p\}_{p \in \mathbb{N}} \) in \( E \):

(i) if \( \{x_p\}_{p \in \mathbb{N}} \) is inf-computable, then its limit in the lower bundle \( \cup_p x_p \) is inf-computable;

(ii) if \( \{x_p\}_{p \in \mathbb{N}} \) is sup-computable, then its limit in the upper bundle \( \cap_p x_p \) is sup-computable.

Proof:

(i) We know that inf-computability amounts to recursive enumerability.
Let $\psi : \mathbb{N}^2 \rightarrow \mathbb{N}$ be the function associated with the sequence $\{x_p\}_{p \in \mathbb{N}}$. The function:

$$u : \mathbb{N}^2 \rightarrow \mathbb{N}, (m, n) \rightarrow \frac{1}{2} (n+m)(n+m+1) + m,$$

is primitive recursive and bijective. Its inverse:

$$v : \mathbb{N} \rightarrow \mathbb{N}^2, \quad p \rightarrow (p_1, p_2);$$

is also primitive recursive and it enumerates $\mathbb{N}^2$ along the "little diagonals", going from left to right:

$$(0, 0), (1, 0), (0, 1), (2, 0), (1, 1), (0, 2), (3, 0), \ldots$$

Therefore:

$$g(p) = \psi(p_1, p_2) = \psi(U_{1/2}^1(v(p)), U_{1/2}^2(v(p))),$$

is a recursive enumeration of the (indices of the) spectrum of $U_p$. Hence $\{n : a_n \subseteq \bigcup_p x_p\}$ is r.e., i.e., $\bigcup_p x_p$ is inf-computable.

(ii) Similarly, we have:

$$g(p) = \psi(p_1, p_2) = \psi(U_{1/2}^1(v(p)), U_{1/2}^2(v(p))),$$

is a recursive enumeration of the spectrum of $\bigcap_p x_p$. Therefore $\{m : b_m \supseteq \bigcap_p x_p\}$ is recursively enumerable, i.e., $\bigcap_p x_p$ is sup-computable. $\Box$

2.3. C-domains

Let $X$ be a poset. An element $u \in X$ is algebraic iff for any decreasing (computable) chain $\{x_k\}_{k \in \mathbb{N}}$ which has a glb $\prod_k x_k$ in $X$:

$$u \supseteq \prod_k x_k \Rightarrow \exists k \quad u \supseteq x_k.$$

As an example, in the poset $E$ ordered by inclusion, every $b_m$ is algebraic. More generally, in any c. p. o. for the opposite order which has an enumerable set of compact elements, every compact element is algebraic.

**Definition:** C-domains. A poset $X$ is a c-domain iff:

(i) $X$ has a largest element.
(ii) Every decreasing computable sequence has a greatest lower bound in $X$.

(iii) An enumeration $a : \mathbb{N} \to X$ of the set of algebraic elements is given, and this enumeration verifies:

$$\forall x \in X, \ x = \bigcap \{ a_n : a_n \geq x \}$$

and $\{ n : a_n \geq x \}$ is recursively enumerable. □

Examples of $c$-domains included in $E$ are

$$X = \{ \{ n \} : n \in \mathbb{N} \} \cup \{ \mathbb{N} \}$$

and

$$X = Y \cup \{ \mathbb{N} \},$$

where $Y \subseteq \mathcal{P}_{\text{fin}}(\mathbb{N})$ is any set of finite subsets of $\mathbb{N}$ which is closed under finite intersection. However some $c$-domain included in $E$ plays a special role:

**Lemma 4:** The set $E_s$ of sup-computable elements of $E$ is the unique $c$-domain such that:

(i) It is included in $E = \mathcal{P}(\mathbb{N})$ as a poset.

(ii) It has the enumeration $b$ and contains every recursive subset $x \subseteq \mathbb{N}$ as an element. □

**Proof:**

(i) $E_s$ is a $c$-domain by Lemma 3.

(ii) Let $X \subseteq \mathcal{P}(\mathbb{N})$ be a $c$-domain containing every computable (recursive) element of $E$. Then $b_m \in X$ for every $m$, whence every sup-computable element is in $X$. Thus $E_s \subseteq X$. If $x \not\in E_s$, then $\{ m : b_m \geq x \}$ is not r.e., i.e., $x \not\in X$. Therefore $X = E_s$. □

In fact it appears that, once $b : \mathbb{N} \to E$ is chosen, $E_s$ is, by construction, the largest $c$-domain contained in $E = \mathcal{P}(\mathbb{N})$.

3. COMPUTABLE FUNCTIONS

3.1. Computable functions

A function $f : E \to E$ is said to be regular iff it is regular for the lower bundle i.e., iff:

$$\forall x \in E, \ f(x) = \bigcup_{a \leq x} f(a) = \bigcup_{n \in x} f(\{ n \}).$$

A regular function $f : E \to E$ is computable iff the set $\{ (m, n) : f(a_n) \leq b_m \}$ is recursively enumerable in the indices $m, n$. 

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Since \( f = \Pi \{ [a_m, b_m] : f(a_n) \leq b_m \} \), the definition of a computable function amounts to the definition of a recursive enumeration of the set of threshold functions:

\[
\{ [a_m, b_m] : f(a_n) \leq b_m \};
\]

approximating \( f \). If we define:

\[
(E \rightarrow E) = \{ f : E \rightarrow E \mid f \text{ regular} \};
\]

\[
[E \rightarrow E] = \{ f : E \rightarrow E \mid f \text{ computable} \};
\]

supplied with the extensional order:

\[
f \leq g \iff \forall x \ f(x) \leq g(x);
\]

there is a canonical bijection between \([E \rightarrow E]\) [resp. \((E \rightarrow E)\)] and the set of principal lowersets of \([E \rightarrow E]\) [resp. of \((E \rightarrow E)\)]:

\[
\downarrow [E \rightarrow E] = \{ \downarrow f : f \in [E \rightarrow E] \};
\]

where \( \downarrow f = \{ g \in [E \rightarrow E] : g \leq f \} \) (resp. \( \ldots \), defined by:

\[
[E \rightarrow E] \leftrightarrow \downarrow [E \rightarrow E];
\]

\[
s(y) = \bigsqcup \{ y = \max y ;
\]

\[
r(x) = \downarrow x = \{ z : z \leq x \}.
\]

Thus the above definition of computability gives a better insight in Nolin's definition of an algorithm (= a principal lowerset) through intersections:

\[
A = \bigcap_{i \in I} FX_i Y_i;
\]

where \( X_i = \downarrow x_i, \ Y_i = \downarrow y_i, \ FX_i Y_i = \downarrow [x_i, y_i] \). Here what is requested by the definition is a recursive enumeration of the family \( \{ FX_i Y_i \}_{i \in I} \) i.e., \( \{ [x_i, y_i] \}_{i \in I} \). This is made precise by the following lemma.

**Lemma 5:** The threshold function:

\[
[x, y] = \lambda t \in E \text{ if } t \leq x \text{ then } y \text{ else } \mathbb{N};
\]

is computable if and only if \( x \) is inf-computable and \( y \) is sup-computable.

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Proof:
\[
\{(m, n) : [x, y] (a_n \subseteq b_m)\} \\
= \{(m, n) : a_n \subseteq x \text{ then } y \text{ else } \mathbb{N} \subseteq b_m\} \\
= \{(m, n) : a_n \subseteq x \text{ and } y \subseteq b_m\} \cup \{(m, n) : a_n \nsubseteq x, \mathbb{N} = b_m\} \\
= \{m : y \subseteq b_m\} \times \{n : a_n \subseteq x\} \cup \{n : a_n \nsubseteq x\}.
\]

Hence \(\{(m, n) : [x, y] (a_n \subseteq b_m)\}\) is r.e. iff \(\{m : y \subseteq b_m\}\) is r.e. and \(\{n : a_n \subseteq x\}\) is recursive [1]. Whence the lemma by Lemma 1. \(\Box\)

It is worthwhile to notice here that computable functions \(f : E \to E\) may be coded, via the bijection \((u, v)\) as r.e. subsets of \(\mathbb{N}\), i.e., \(\forall f : E \to E\) computable we define:

\[
\text{graph}^*(f) = \{(m, n) | f(a_n) \subseteq b_m\} = \{(m, n) | f(\{n-1\}) \subseteq b_m\}.
\]

The corresponding definition in \(\mathcal{P} \omega [8]\) is:

\[
\text{graph}(f) = \{(m, n) | \{m\} \subseteq f(e_n)\}, \quad e_n = \emptyset b_n.
\]

Any element \(u \in \mathcal{P} \omega\) operates as a continuous function by:

\[
\text{fun}(u)(x) = \{m : \exists e_n \subseteq x : (n, m) \in u\}.
\]

The corresponding definition for our functions is for any \(u \in E = \mathcal{P}(\mathbb{N})\):

\[
\text{fun}^*(u)(x) = \bigcup_{n \in x} (\{b_m : (n+1, m) \in u\}).
\]

One easily verifies that:

(i) for any regular \(f : E \to E\), \(f = \text{fun}^*(\text{graph}^*(f))\);
(ii) for any \(u \in E\), \(u \subseteq \text{graph}^*(\text{fun}^*(u))\), where the equality holds iff:

\[
\{(p, q) : \cap \{b_m : (p+1, m) \in u\} \subseteq b_q\} \subseteq u;
\]

(iii) Any computable \(f : E \to E\) yields, by definition, a recursively enumerable graph\(^*\)(\(f\)). Conversely, a r.e. set \(u \in E\) defines a computable function \(\text{fun}^*(u) : E \to E\) iff the set:

\[
\{(p, q) : \cap \{b_m : (p+1, m) \in u\} \subseteq b_q\} \\
= \{(p, q) : \cup \{e_m : (p+1, m) \in u\} \nsubseteq e_q\} \\
= \{(p, q) : \exists m : (p+1, m) \in u \land e_q \nsubseteq e_m\};
\]

is recursively enumerable.

The relation \(e_q \nsubseteq e_m\) is recursive, therefore the set is recursively enumerable; i.e. every r.e. set \(u \in E\) defines a computable function \(\text{fun}^*(u) : E \to E\).

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However, this analogy between $P$ and our construction will become fuzzier at higher levels of functionality, due to our use of Wadsworth scheme (cf. infra).

Actually we can define $F_{ab} = \downarrow \{a, b\}$:

$$\downarrow \{a, b\} = \{f \in [E \to E] : f(a) \leq b\} = \{f \in [E \to E] : f \leq [a, b]\}.$$  

Obviously $\forall f \in [E \to E]$:

$$\downarrow f = \bigcap \{\downarrow [a, b] : f \leq [a, b]\};$$

or:

$$\uparrow f = \bigcup \{\uparrow [a, b] : f(a) \leq b\}.$$  

The relation:

$$\forall f \in (E \to E), \ f = \Pi \{[a_n, b_m] : f(a_n) \leq b_m\};$$

gives a bundle structure over the set of regular functions $(E \to E)$; more precisely this makes $(E \to E)$ an elementary monic ordered bundle, with:

$$s(f) = \{[a_n, b_m] : f(a_n) \leq b_m\};$$

as a spectrum function. This structure may be made algebraic by taking the closure of the spectra for finite greatest lower bounds. The set $[E \to E]$ forms a subbundle for this structure.

An enumeration of the kernel is given by:

$$\beta : m \rightarrow \beta_m = e_{k_0} \Pi \ldots \Pi e_{k_{p-1}};$$

with:

$$m = \sum_{i < p} 2^i \ k_0 < k_1 < \ldots < k_{p-1};$$

$$e_{k_i} = [a_{u_1}\omega(k_i), b_{u_2}\omega(k_i)];$$

where function $\omega$, here applied to $k_i$, is the inverse of function $u$ (recursive enumeration of $\mathbb{N}^2$, cf. supra).

The set of $\beta_m$'s is closed under finite $\Pi$ and every $\beta_m$ verifies the following algebraicity property:

$$\forall \{x_i\}_{i \in I}, \ \beta_m \geq \prod_{i \in I} x_i \Rightarrow \exists \beta_{m_i} \geq x_i.$$  

Moreover, for the computability aspect, we have.
LEMMA 6: The relations $\beta_m \leq \beta_n$, $[a_n, b_m] \geq \beta_p$, $\beta_m = \beta_n \Pi \beta_p$ are all recursive in the indices $m, n, p$.

Proof: We consider the proof for $\beta_m \leq \beta_n$. First notice that:

$$[a, b] \leq [c, d] \iff (c \leq a \land b \leq d) \quad \text{or} \quad d = \mathbb{N}.$$ 

Thus:

$$[a_n, b_m] \leq [a_n', b_{m'}] \iff a_n' \leq a_n \text{ (recursive)} \land b_m \leq b_{m'} \text{ (recursive)};$$

or:

$$b_{m'} = \mathbb{N} \text{ (recursive).}$$

Therefore the relation $[a_n, b_m] \leq [a_n', b_{m'}]$ is recursive in the indices $n, m, n', m'$. Now consider $\beta_m \leq \beta_n$. We have:

$$\beta_m = e_{k_0} \Pi \ldots \Pi e_{k_{p-1}}, \quad m = \sum_{i < p} 2^{k_i}, \quad e_{k_i} = [a_{v_i}(\nu(k_i)), b_{v_i}(\nu(k_i))];$$

$$\beta_n = e_{l_0} \Pi \ldots \Pi e_{l_{q-1}}, \quad n = \sum_{j < q} 2^{l_j}, \quad e_{l_j} = [a_{u_j}(\nu(l_j)), b_{u_j}(\nu(l_j))].$$

Thus:

$$\beta_m \leq \beta_n \iff e_{k_0} \Pi \ldots \Pi e_{k_{p-1}} \leq e_{l_0} \Pi \ldots \Pi e_{l_{q-1}}$$

$$(\text{simplifying the notations})$$

$$[a_{k_0}, b_{k_0}] \Pi \ldots \Pi [a_{k_{p-1}}, b_{k_{p-1}}] \leq [a_{l_0}, b_{l_0}] \Pi \ldots \Pi [a_{l_{q-1}}, b_{l_{q-1}}]$$

$$\iff \bigcap \{b_{k_i} : a_{l_0} \leq a_{k_i}\} \subseteq \bigcap \{b_{l_j} : a_{l_0} \leq a_{l_j}\}$$

and

$$\bigcap \{b_{k_i} : a_{l_1} \leq a_{k_i}\} \subseteq \bigcap \{b_{l_j} : a_{l_1} \leq a_{l_j}\};$$

and

$$\bigcap \{b_{k_i} : a_{l_{q-1}} \leq a_{k_i}\} \subseteq \bigcap \{b_{l_j} : a_{l_{q-1}} \leq a_{l_j}\}.$$

Each of these inequalities is decidable. Thus the inequality $\beta_m \leq \beta_n$ is decidable (recursive) in the indices $m, n$. 

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We have a similar argument for the other relations. Whence the lemma.

**LEMMA 7:** The set of computable functions $f : E \rightarrow E$ which have the following simplicity property:

$$\forall a_n, \exists m, \quad f(a_n) = a_m;$$

is closed under composition.

**Proof:** Let $E \xrightarrow{f} E \xrightarrow{g} E$ be two computable functions.

- $f$ computable $\iff \{(m, n) : [a_m, b_m] \subseteq f\}$ is r.e.;
- $g$ computable $\iff \{(p, q) : [a_p, b_q] \subseteq g\}$ is r.e..

The spectrum of $g \circ f$, which is a monotone function, is given by:

$$s(g \circ f) = \{[a_n, b_n] : \exists m [a_m, b_m] \in s(f) \quad \exists p, a_p \subseteq b_m \text{ and } [a_p, b_q] \in s(g)\};$$

where $x \subseteq a_n \Rightarrow f(x) \subseteq b_m$ and $g(b_m) = \bigcup \{g(a_p) : a_p \subseteq b_m\}$.

The relation $a_p \subseteq b_m$ is recursive, thus it is enough to transform through $g$ the elements $a_p \subseteq b_m$ in order to find out $s(g \circ f)$. If $b_m = \bigcup_i a_{p_i}$, then we have the diagram:

$$a_{p_0} \xrightarrow{g} b_{q_0};$$
$$a_{p_1} \rightarrow b_{q_1};$$
$$a_n \rightarrow b_m = \vdots$$
$$a_{p_i} \rightarrow b_{q_i};$$
$$\vdots$$

This gives $[a_n, \bigcup_i b_{q_i}]$ as approximating $g \circ f$, i.e., $(g \circ f)(a_n) \subseteq \bigcup_i b_{q_i}$.

The set $b_m = \bigcup_i a_{p_i}$ is recursive, thus the sequence $\{b_{q_i}\}$ is recursively enumerable since $\{(p, q) : [a_p, b_q] \subseteq g\}$ is recursively enumerable.

Therefore $\bigcup_i b_{q_i}$ is r.e., i.e., *inf-computable*. What we need is the sup-computability of $\bigcup_i b_{q_i}$ in order to have a r.e. decomposition:

$$[a_n, \bigcup_i b_{q_i}] = \prod_k [a_n, b_k],$$

from which we would deduce a r.e. spectrum for $g \circ f$. But if we impose that $g \circ f$ be simple, i.e., $f(a_n) = a_m$ for some $m \in \mathbb{N}$, then, since:

$$g = \prod \{[a_p, b_q] : g(a_p) \subseteq b_q\},$$

then:

$$(g \circ f)(a_n) = g(a_m) = \bigcap_q \{b_q : a_m \subseteq a_p\}$$

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which is a sup-computable element of $E$, whence an r. e. spectrum for $g \circ f$:

$$s(g \circ f) = \{ [a_n, b_s] : f(a_n) = a_p, g(a_p) \leq b_s \} .$$

**Lemma 8:** There is a canonical bijection between the simple computable functions $f : E \to E$ and the recursive functions from $\mathbb{N}$ to $\mathbb{N}$.

**Proof:** Obvious. □

**Lemma 9:** Let $f : E \to E$ be a computable function and $\{ x_p \}_{p \in \mathbb{N}}$ be an inf-computable sequence of rational elements of $E$. Then the image $\{ f(x_p) \}_{p \in \mathbb{N}}$ of the sequence is a sup-computable sequence of $E$.

**Proof:** $(x_p)_{p \in \mathbb{N}}$ infinit-computable sequence $\iff \exists \psi : \mathbb{N}^2 \to \mathbb{N}$ recursive such that $\psi(p, .)$ enumerates $s(x_p)$. $f$ computable $\iff \{ (m, n) : f(a_n) \leq b_m \}$ is r. e. Since every $x_p$ is rational, $f(x_p) = \cap \{ b_m : x_p \leq a_m, f(a_n) \leq b_m \}$ the relation $x_p \leq a_n$ is recursive, the relation $f(a_n) \leq b_m$ is recursively enumerable, thus the set $\{ m : x_p \leq a_n, f(a_n) \leq b_m \}$ is r. e. which implies that $\{ m : b_m \geq f(x_p) \}$ is r. e. Whence a recursive function $\psi : \mathbb{N}^2 \to \mathbb{N}$ such that $\psi(p, .)$ enumerates the spectrum of $f(x_p)$ for the upper bundle:

$$\psi(p, q) = [\text{enumeration of } \{ m : b_m \geq f(x_p) \}](q).$$

Thus $\{ f(x_p) \}_{p \in \mathbb{N}}$ is sup-computable. □

### 3.2. Computable sequences of functions

**Definition:** A sequence $\{ f_p \}_{p \in \mathbb{N}}$ of regular functions from $E$ to $E$ is a computable sequence iff $\exists \psi : \mathbb{N}^2 \to \mathbb{N}$ recursive such that $\forall p \in \mathbb{N}, \psi(p, .)$ is an enumeration of the spectrum of $f_p$. □

In particular every $f_p$ will be a computable function. We have an analogous of Lemma 3 (i) for computable sequences of functions:

**Lemma 10:** If $\{ f_p \}_{p \in \mathbb{N}}$ is a computable sequence of functions, then its greatest lower bound $\prod_p f_p$ is a computable function.

**Proof:** The function $\prod_p f_p$ is regular since $(\prod_p f_p)(a_n) = \prod_p f(a_p)$ for every rational $a_m$ and we take the least regular extension of this to non rational elements:

$$(\prod_p f_p)(x) = \cup \{ (\prod_p f_p)(a_n) : a_n \leq x \} .$$
Now we just have to glue the spectra together:

\[ s(\prod f_p) = \bigcup_p s(f_p). \]

By means of the indices, the sequence \( \{ s(f_p) \}_{p \in \mathbb{N}} \) defines an inf-computable sequence of \( E \). Hence \( \bigcup_p s(f_p) \) is inf-computable, therefore recursively enumerable.

Let us define for any regular \( f \in (E \rightarrow E) \), \( f \) is finitely computable iff the set \( \{(m, n) : f(a_n) \leq b_n\} \) is recursive. Then we have the analogous of Lemma 4:

**Lemma 11**: Given the enumeration \( \beta \) of the algebraic elements, the set of computable functions \( [E \rightarrow E] \) is the unique c-domain such that:

(i) it is included in \((E \rightarrow E)\) as a poset;

(ii) it contains every finitely computable function.

**Proof**:

(i) \([E \rightarrow E]\) is a c-domain: The largest element of \([E \rightarrow E]\) is the constant function \( x \rightarrow \mathbb{N} \). Every decreasing computable sequence has a glb in \([E \rightarrow E]\) by Lemma 10. The other requirements are trivially fulfilled.

(ii) Let \( X \subseteq (E \rightarrow E) \) be a c-domain containing the finitely computable functions. Then every \( \beta_m \) is in \( X \), therefore \( E_s \) which is the closure of \( \{ \beta_m : m \in \mathbb{N} \} \) for the glb’s of decreasing computable sequence is included in \( X \). Thus \([E \rightarrow E] \subseteq X \) and one easily sees that \([E \rightarrow E] = X\).

We obtain similar results by replacing \( E \) by \( E_s \), and there is a canonical bijection between \([E \rightarrow E]\) and \([E_s \rightarrow E_s]\).

4. WADSWORTH SCHEME

We have a c-domain structure over \([E \rightarrow E] = \Delta_1\). We can define the partial computable function spaces \([E \rightarrow \Delta_1], [\Delta_1 \rightarrow E], [\Delta_1 \rightarrow \Delta_1]\):

\([E \rightarrow \Delta_1]\) = \{ \( f : f(x) = \bigcup_{a_n} f(a_n) \) and \( \{(m, n) : f(a_n) \leq \beta_m\} \) is r.e. \};

\([\Delta_1 \rightarrow E]\) = \{ \( f : f(\prod_k x_k) = \prod_k f(x_k) \) for every decreasing computable sequence \( (x_k) \) and \( \{(m, n) : f(\beta_n) \leq b_n\} \) is r.e. \};

\([\Delta_1 \rightarrow \Delta_1]\) = \{ \( f : f(\prod_k x_k) = \prod_k f(x_k) \) for every decreasing computable sequence \( (x_k) \) and \( \{(m, n) : f(\beta_n) \leq \beta_m\} \) is r.e. \}.

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We must be careful here, we are dealing with partial functions in the set-theoretical sense.

**Lemma 12:** The spaces \([E \rightarrow \Delta_1], [\Delta_1 \rightarrow E], [\Delta_1 \rightarrow \Delta_1]\), when supplied with the extensional order, all have a c-domain structure.

**Proof:** Analogous to Lemma 11. Only the rational elements change, and we use the same enumeration technique as for \([E \rightarrow E]\).

Then we can define:

\[
\begin{align*}
\Delta_1 &= [E \rightarrow E], \\
A_1 &= E + \Delta_1; \\
A_2 &= E + \Delta_2 = E + [E + \Delta_1 \rightarrow E + \Delta_1];
\end{align*}
\]

where:

\[
\begin{align*}
\Delta_2 &= [E + \Delta_1 \rightarrow E + \Delta_1] = [E \rightarrow E + \Delta_1] \times [\Delta_1 \rightarrow E + \Delta_1]; \\
\end{align*}
\]

with:

\[
\begin{align*}
[E \rightarrow E + \Delta_1] &= [E \rightarrow E] + [E \rightarrow \Delta_1]; \\
[\Delta_1 \rightarrow E + \Delta_1] &= [\Delta_1 \rightarrow E] + [\Delta_1 \rightarrow \Delta_1]; \\
[E \rightarrow E] + [E \rightarrow \Delta_1] &= \{ f + g : f \in [E \rightarrow E], g \in [E \rightarrow \Delta_1] \}.
\end{align*}
\]

\(f + g\) being defined as the canonical extension of \(f\) and \(g\), if \(\{ \text{Dom} (f), \text{Dom} (g) \}\) is a partition of \(E\). (Here, as has been said earlier, we use partial functions.) The space \([\Delta_1 \rightarrow E] + [\Delta_1 \rightarrow \Delta]\) is defined in a similar way.

This defines the partial sequence of spaces:

\[
\begin{align*}
A_0 &= E_s; \\
A_1 &= E_s + \Delta_1 = E_s + [A_0 \rightarrow A_0]; \\
A_2 &= E_s + \Delta_2 = E_s + [A_1 \rightarrow A_1].
\end{align*}
\]

As may be seen from the construction, \(\Delta_2\) has a c-domain structure by using Lemma 12. We also have the following property: if \(\{ f_p \}_{p \in \mathbb{N}}\) is a computable sequence of \(\Delta_1\) and if \(G : \Delta_1 \rightarrow E + \Delta_1\) is a computable function, then \(\{ G(f_p) \}_{p \in \mathbb{N}}\) is a computable sequence, by an argument similar to the one used for Lemma 9.

The finite sequence of (1) can be extended by:

\[
\begin{align*}
A_0 &= E_s; \\
A_{n+1} &= E_s + \Delta_{n+1} = E_s + [A_n \rightarrow A_n],
\end{align*}
\]
where for any \( n \in \mathbb{N} \), \( \Delta_{n+1} = [A_n \to A_n] \) has a c-domain structure by construction, and \( A_{n+1} \) is supplied with a bundle structure obtained by gluing together the \( E_s \) structure and the \( \Delta_{n+1} \) structure as follows:

\[
\begin{array}{c}
T \\
\downarrow \\
E_s \\
\Delta_{n+1}
\end{array} = A_{n+1},
\]

Every \( A_n \) is a c-domain. Now the threshold function:

\[
[x, y] : z \to \begin{cases} 
1 & \text{if } z \leq x \\
0 & \text{otherwise}
\end{cases}
\]

is computable iff \( x \) is finitely *computable* and \( y \) is sup-computable. If \( x \in E \), this can be checked at once (finite computability = recursiveness). If \( x \in \Delta_n \cup \{ T \} \), we use the definition.

We now make the above sequence a diagram, by defining, following Wadsworth 71:

\[
\begin{align*}
\forall n \in \mathbb{N} ; \\
i_0 &: A_0 \to A_1, & x &\mapsto x \\
i_n &: A_n \to A_{n+1}, & x &\mapsto x & x \in E_s \cup \{ T \} ; \\
\end{align*}
\]

\[
\begin{align*}
i_{n-1} \circ x \circ j_{n-1} &\quad \text{if } x \in \Delta_n ; \\
j_0 &: A_1 \to A_0, & y &\mapsto y & y \in E_s \cup \{ T \} ; \\
j_n &: A_{n+1} \to A_n, & y &\mapsto y & y \in E_s ; \\
j_{n-1} \circ y \circ i_{n-1} &\quad \text{if } y \in \Delta_{n+1} ;
\end{align*}
\]

This diagram will be called *Wadsworth scheme*. Notice that, for every \( n \):

\[
\begin{align*}
j_n \circ i_n &\geq \text{id}_{A_n} ; \\
i_n \circ j_n &\geq \text{id}_{A_n} ;
\end{align*}
\]

\( i_n \) and \( j_n \) are distributive with respect to \( \Pi \) and \( \cup \).

An element \( x = (x_n)_{n \in \mathbb{N}} \in \bigtimes_{n \in \mathbb{N}} A_n \) belonging to the cartesian product of the \( A_n \)'s will be called *computable* if and only if there exists \( \psi : \mathbb{N}^2 \to \mathbb{N} \) recursive such that for every \( n \in \mathbb{N} \), \( \psi(n, .) \) is an enumeration of (the indices of) the spectrum of \( x_n \). If \( x_n \in E \), then we consider the spectrum of \( x_n \) for the upper
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bundle. Computable sequences of elements of $\bigtimes_{n \in \mathbb{N}} A_n$ are defined in the usual way. Notice that $\bigtimes_{n \in \mathbb{N}} A_n$ has a c-domain structure. Its kernel is the cartesian product of the kernels. Consider now the projective limit:

$$A_\infty = \left\{ (x_n)_{n \in \mathbb{N}} \in \bigtimes_{n \in \mathbb{N}} A_n \mid x_n = j_n(x_{n+1}) \right\}.$$ 

Then for any $\{(x_n)_{n \in \mathbb{N}} \in A_\infty\}$, we have:

- either $x_n$ belongs to the E-part of $A_\infty$ and $x_n = x_{n+1}$;
- or $x_n$ belongs to the functional part of $A_\infty$ and $x_n = j_{n-1} \circ x_{n+1} \circ i_n^{-1}$.

Therefore we have the equality of sets:

$$A_\infty = \left\{ (x_n)_{n \in \mathbb{N}} \in \bigtimes_{n \in \mathbb{N}} A_n \mid x_n = 0 \in E_\emptyset \right\} + \left\{ (x_n)_{n \in \mathbb{N}} \in \bigtimes_{n \in \mathbb{N}} A_n \mid x_0 = \mathbb{N}, x_n = j_n(x_{n+1}) \right\}.$$ 

Thus $A_\infty = E_\emptyset + \Delta_\infty$, where $\Delta_\infty$ is the functional part of $A_\infty$.

Now both notions of computability and projective limit are put together in order to define the set of computable projective sequences:

$$A_\omega = \left\{ (x_n)_{n \in \mathbb{N}} \in A_\infty \mid (x_n)_{n \in \mathbb{N}} \text{ is computable} \right\}.$$ 

One easily sees that $A_\omega = E_\emptyset + \Delta_\omega$. The closure of $A_\omega$ for decreasing computable sequences will be called an “algorithm space”. Here again, we disregard the lower bundle structure of $A_\omega$. As far as functional elements are concerned, because of the triviality of this bundle structure. Thus computability here means computability for the upper bundle structure.

DEFINITION: The algorithm space $\mathcal{A}$ is the smallest set containing $A_\omega$ and such that every decreasing computable sequence in $\Delta_\omega$ has a greater lower bound. Elements of $\mathcal{A}$ are called algorithms.

In other words, $\mathcal{A}$ is the smallest c-domain containing $A_\omega$. Its kernel is canonically isomorphic to the union of the kernels of the $A_n$'s:

$$N(\mathcal{A}) = \bigcup_n N(A_n).$$

LEMMA 13: Let $i_{n\omega} : A_n \to \mathcal{A}$ be the canonical injection of $A_n$ into $\mathcal{A}$, and $j_{n\omega} : \mathcal{A} \to A_n$ the canonical projection of $\mathcal{A}$ onto $A_n$. Then both $i_{n\omega}$ and $j_{n\omega}$ are computable.

Proof: The regularity comes from the distributivity of functions $i$ and $j$. 

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(i) Injection \( i_{\infty} \): this set must be r. e.:

\[
\{(p, q) : i_{\infty}(a_p) \leq b_q \} = (\text{if } b_q = (b^q_n)_{n \in \mathbb{N}}) \{(p, q) : a_p \leq b^q_q \text{ in } A_n; \ b_q = i_{\infty}(b^q_q)\}.
\]

We know that in \( A_n \) the relation \( a_p \leq b^q_q \) is recursive, and the set of rational elements of \( A_n \) is enumerable (this set contains all the rationals of \( E_q \) for the upper bundle, and all the rationals of \( \Delta_n \)), which completes the proof.

(ii) Projection \( j_{\infty} \): Similarly this set must be r. e.:

\[
\{(p, q) : i_{\infty}(a_p) \leq b_q \} = (\text{if } a_p = (a^p_n)_{n \in \mathbb{N}}) = \{(p, q) : a^p_p \leq b^p_q \} = \{(p, q) : a^p_p \leq b^q_q \};
\]

which is r. e. since the relation \( a^p_p \leq b^q_q \) is recursive in \( A_p \) by the same argument as for Lemma 6. Whence the lemma.

**Lemma 14:** Let \( f : \mathcal{A} \rightarrow \mathcal{A} \) be a computable function. Then the sequence of \( X A_n \) defined by:

\[
\{ f[p]_{p \in \mathbb{N}} \}
\]

is projective and computable, and thus an element of \( X A_\omega \). □

**Proof:**

(i) The sequence \( \{ f[p]_{p \in \mathbb{N}} \} \) is projective since:

\[
j_n(\{ f[n+1] \}) = j_{n-1} \circ f[n+1] \circ i_{n-1}
\]

\[
= (j_{n-1} \circ j_\omega) \circ f \circ (i_{\omega} \circ i_{n-1})
\]

\[
= j_\omega \circ i_{n-1} \circ f \circ i_{n-1} = \{ f[p] \}.
\]

Thus \( \{ f[p]_{p \in \mathbb{N}} \} \in A_\omega. \)

(ii) The sequence \( \{ f[p]_{p \in \mathbb{N}} \} \) is computable:

spectrum \( \{ f[0] \} = \mathbb{N}, \)

spectrum \( \{ f[n+1] \} = j_\omega \circ \text{spectrum } f \circ i_{\omega}. \)

Let \( \psi \) be a function defined as follows:

1. \( \psi(0, .) = \lambda \cdot q. \) index of \( \mathbb{N} \) in the enumerating of the kernel of \( A_\omega; \)
2. \( \psi(p, q) = j_\omega \circ i_{p-1} \circ s(q) \circ i_{p-1} \) where \( s : q \rightarrow s(q) \) is a recursive enumeration of the spectrum of \( f. \)

Thus \( \psi : \mathbb{N}^2 \rightarrow \mathbb{N} \) seen as a function into the indices is recursive and \( \psi(p, .) \) is an enumeration of the spectrum of \( \{ f[p] \}. \) Thus \( \{ f[p]_{p \in \mathbb{N}} \} \) is a computable element of \( X A_\omega. \) Therefore:

\[
= \{ f[p]_{p \in \mathbb{N}} \} \in A_\omega. \ □
\]

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LEMMA 15: Any $x \in \Delta_\omega$ defines a computable function:

$$[x] : \mathcal{A} \rightarrow \mathcal{A};$$

$$y \rightarrow \prod_{n} x_{n+1} (y_n) \quad \text{if} \quad y \in \Delta_\omega \quad \text{or} \quad y = a_p \in E;$$

$$\cup \{ [x] (a_p) : a_p \subseteq y \in E \} \text{ otherwise.}$$

Proof: Regularity here means regularity for the c-domain structure, and is obvious. The spectrum of $[x]$ is the cartesian product $\prod_n s(x_n)$, which is r.e. since sequence $x = (x_n)_{n \in \mathbb{N}}$ is computable.

More importantly, we have a representation property, which justifies the use of the c-domain notion:

LEMMA 16: Let $f : \mathcal{A} \rightarrow \mathcal{A}$ be a computable function. Then for any $z$ in the kernel of $\Delta_\omega$ or such that $z = a_p \in E$, the following are equivalent:

(i) $f(z) = [f][z]$;

(ii) if $z = (z_n)_{n \in \mathbb{N}} = \prod_n z_n$, then $f(z) = f(\prod z_n) = \prod_n f(z_n)$.

Proof: It suffices to see that:

$$[f][z] = \prod_n [f][n+1] (z_n) = \prod_n (\lambda y \in \Delta_\omega. (f(y))_n) (z_n) = \prod_n (f(z_n))_n$$

$$= \prod_{n,k} (f(z_n))_k = \text{(property of \Pi)} = \prod_{n,k} (f_\Pi (z_n))_k = \prod f(z_n).$$

Thus $[f][z] = f(z)$ is equivalent to $\prod f(z_n) = f(z)$,

Whence the lemma.

Lemma 16 characterizes completely the representation of computable functions as elements of $\Delta_\omega$, because of the definition of the operation $[.] : \Delta_\omega \rightarrow [\mathcal{A} \rightarrow \mathcal{A}]$ and since $\mathcal{A}$ is a c-domain.

Thus, basically, we see that our functions must satisfy the following "continuity" property:

$$\forall \{ x_k \}_{k \in \mathbb{N}} \text{ decreasing computable sequence} \quad f(\prod x_k) = \prod f(x_k);$$

which is analogous to Scott-continuity, but for two modifications: the inversion of the order and the introduction of computability. In the present setting, the above property is what we need when computing with procedures.
THEOREM (existence of enumerable algorithm domains): There exists an enumerable set $\mathcal{A}$, called algorithm space, such that:

(i) $\mathcal{A} = E_s + \mathcal{F}$, and every element is computable.
(ii) Every decreasing computable sequence has a greatest lower bound in $\mathcal{A}$.
(iii) if $x, y \in \mathcal{A}$ then any computable threshold function:

$$[x, y] : z \rightarrow \text{if } z \leq x \text{ then } y \text{ else } T;$$

is represented as an element of $\mathcal{F}$ which is:

$$(\lambda z \in A_n. \prod_p \text{ if } z_p \leq x_p \text{ then } y_n \text{ else } T_n)_{n+1} \in \mathbb{N}.$$ 

Proof: Results from the preceding lemmae.

The enumerability of $\mathcal{A}$ comes from the fact that the set of recursive functions $\Psi : \mathbb{N}^2 \rightarrow \mathbb{N}$ is enumerable, and all our objects are computable. □

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