On the Complexity of Inductive Definitions

DOUGLAS CENZER\textsuperscript{1} and JEFFREY B. REMMEL\textsuperscript{2}\textsuperscript{†}

\textsuperscript{1} Department of Mathematics, University of Florida, P.O. Box 118105, Gainesville, Florida 32611, USA.

\textsuperscript{2} Department of Mathematics, University of California at San Diego, La Jolla, CA 92093, USA.

Received September 2005

We study the complexity of computable and $\Sigma^0_1$ inductive definitions of sets of natural numbers. For example, we show how to assign natural indices to monotone $\Sigma^0_1$-definitions and we use these to calculate the complexity of the set of all indices of monotone $\Sigma^0_1$-definitions which are computable. We also examine the complexity of a new type of inductive definition which we call \textit{weakly finitary} monotone inductive definitions. Applications are given in proof theory and in logic programming.

\section{Introduction}

Inductive definitions play a central role in mathematical logic and computer science. For example, the set of formulas in a first order language is given by an inductive definition. Given a set $A$ of axioms for a mathematical theory $T$ and a set of logical axioms and rules, the theory $T$ is obtained by an inductive definition. The set of computable functions can be realized by an inductive definition. Similarly, for any Horn logic program $P$, the unique stable model of $P$ is obtained by an inductive definition.

It is well-known that for any computable or $\Sigma^0_1$ monotone inductive definition $\Gamma$, one can construct the closure of $\Gamma$, $\text{Cl}(\Gamma)$, in at most $\omega$ steps and $\text{Cl}(\Gamma)$ is always a $\Sigma^0_1$ set. In some situations, it is important that $\text{Cl}(\Gamma)$ is computable. For example, it is important that the set of formulas in a typical first order theory is computable. In other situations, we know that $\text{Cl}(\Gamma)$ is $\Sigma^0_1$ but not computable. For example, even a finitely axiomatizable theory $T$ may be $\Sigma^0_1$ but not decidable (computable). In this paper, we shall explore the complexity of various properties of the closure of a $\Sigma^0_1$ monotone inductive definition $\Gamma$. For examples, we shall consider properties like when the closure of $\Gamma$ is finite, cofinite, or computable or when the closure ordinal of $\Gamma$ is finite or equal to $\omega$. We shall do this by assigning indices to $\Sigma^0_1$ monotone inductive operators. In particular, this means that we can effectively enumerate the family of all $\Sigma^0_1$ monotone inductive operators as $\Gamma_0, \Gamma_1, \ldots$.

\textsuperscript{†} This work partially supported by NSF under grant DMS 0400507.
Then, for example, we shall show that the set $C$ of indices $e$ such that the closure or least fixed point $lfp(\Gamma_e)$ is computable is $\Sigma^0_3$ complete.

We will also define a new class of inductive operators called weakly finitary monotone inductive operators. The basic idea is that for a weakly finitary operator there may exist a finite set of elements $x$ such that $x$ is forced into $\Gamma(A)$ only if $A$ contains one of a collection of possibly infinite sets. We will show that if $\Gamma$ is a weakly finitary monotone inductive operator, then it will still be the case that $lfp(\Gamma)$ will be $\Sigma^0_3$ but that it can take more than $\omega$ steps to construct $lfp(\Gamma)$. An example of such an operator is when we allow finitely many instances of the $\omega$-rule to generate a partial theory of arithmetic.

We also assign indices to the family of weakly finite $\Sigma^0_1$ monotone inductive operators. We shall show that the set of indices of weakly finitary $\Sigma^0_1$ monotone inductive operators $\Gamma$ such that $lfp(\Gamma)$ is computable is also $\Sigma^0_3$ complete. However, for certain computable sets $R$, the set of indices of weakly finitary $\Sigma^0_1$ monotone inductive operators $\Gamma$ such that $lfp(\Gamma) \cap R$ is computable lies in the difference hierarchy over the $\Sigma^0_3$ sets.

We will use standard notation from computability theory (Soare 1987). Let $\mathbb{N}$ denote the set of natural numbers and $\mathcal{P}(\mathbb{N})$ denote the set of all subsets of $\mathbb{N}$. In particular, we let $\phi_e (\phi^A_e)$ denote the $e$-th partial computable function ($e$-th $A$-partial computable function) from $\mathbb{N}$ to $\mathbb{N}$ and let $W_e = \text{Dom}(\phi_e) (W^A_e = \text{Dom}(\phi^A_e))$ be the $e$-th computably enumerable (c.e.) ($e$-th $A$-computably enumerable) subset of $\mathbb{N}$. Note that computably enumerable and recursively enumerable (r.e.) have the same meaning and likewise computable functions are also known as recursive functions. We let $W_{e,s} (W^A_{e,s})$ denote the set of numbers $m \leq s$ such that $\phi_e(m) (\phi^A_e(m))$ converges in $s$ or fewer steps. Given a finite set $S = \{a_1 < \ldots a_n\}$, the canonical index of $S$ is $\sum_{i=1}^{n} 2^{a_i}$. The canonical index of the empty set is 0. We let $D_n$ denote the finite set whose canonical index is $n$.

We fix a primitive recursive pairing function, $[x, y] = \frac{1}{2}((x+y)^2 + 3x + y)$ from $\mathbb{N} \times \mathbb{N}$ to $\mathbb{N}$. For any sequence $a_1, \ldots, a_n$ with $n \geq 3$, we define $[a_1, \ldots, a_n]$ by the usual inductive procedure of defining $[a_1, \ldots, a_n] = [a_1, [a_2, \ldots, a_n]]$. The explicit index of the sequence $(a_1, \ldots, a_n)$ is defined by $\langle a_1 \rangle = [1, a_1]$ if $n = 1$ and $\langle a_1, \ldots, a_n \rangle = [n, [a_1, \ldots, a_n]]$ if $n \geq 2$.

## 2. Inductive Definitions

In this paper, we are going to consider inductive operators $\Gamma : \mathcal{P}(\mathbb{N}) \rightarrow \mathcal{P}(\mathbb{N})$ which inductively define subsets of $\mathbb{N}$. We begin with a review of basic definitions and results which can be found, for example, in Hinman (Hinman 1978).

**Definition 2.1.** Let $\Gamma : \mathcal{P}(\mathbb{N}) \rightarrow \mathcal{P}(\mathbb{N})$.

1. $\Gamma$ is said to be **monotone** if $A \subset B$ implies $\Gamma(A) \subseteq \Gamma(B)$ for all $A, B$.
2. $\Gamma$ is said to be **inclusive** if $A \subseteq \Gamma(A)$ for all $A$.
3. $\Gamma$ is said to be **inductive** if it is either monotone or inclusive.

An inductive operator $\Gamma$ recursively defines a sequence \{$_{\alpha}^\Gamma : \alpha$ an ordinal\} by setting $\Gamma^0 = \emptyset$, $\Gamma^{\alpha+1} = \Gamma(\Gamma^\alpha)$ for all $\alpha$ and $\Gamma^\lambda = \bigcup_{\alpha < \lambda} \Gamma^\alpha$. It is easy to see that $\Gamma^\alpha \subseteq \Gamma^\beta$ whenever $\alpha < \beta$. By cardinality considerations, there exists a countable ordinal $\alpha$ such that $\Gamma^\alpha = \Gamma^\beta$ for all $\beta > \alpha$. The least such $\alpha$ is called the **closure ordinal** of $\Gamma$ and will
be denoted by $|\Gamma|$. The set $\Gamma^{|\Gamma|}$ is called the closure of $\Gamma$ or the set inductively defined by $\Gamma$ and will be denoted by $Cl(\Gamma)$.

For a monotone operator, the closure is also the least fixed point $lfp(\Gamma)$ as indicated by the following lemma, see Hinman (Hinman 1978).

**Lemma 2.1.** If $\Gamma$ is a monotone operator, then $Cl(\Gamma)$ is the unique least set $C$ such that $\Gamma(C) = C$. In fact, for any set $A$, $\Gamma(A) \subseteq A$ if and only if $cl(C) \subseteq A$.

For any operator $\Gamma : \mathcal{P}(\mathcal{N}) \to \mathcal{P}(\mathcal{N})$, let $R_\Gamma \subseteq \mathbb{N} \times \mathcal{P}(\mathcal{N})$ be given by $R_\Gamma(m, A) \iff m \in \Gamma(A)$. In general, we say that a predicate $R(x_1, \ldots, x_n, A) \subseteq \mathbb{N}^k \times \mathcal{P}(\mathcal{N})$ is computable if there is an oracle Turing machine $M_e$ such that for any $A \in \mathcal{P}(\mathcal{N})$, $M_e$ with oracle $A$ and input $(x_1, \ldots, x_n)$ outputs 1 if $R(x_1, \ldots, x_n, A)$ holds and outputs 0, otherwise. The notation of a predicate being $\Sigma_0^n$, $\Pi_0^n$, $\Sigma_1^1$, $\Pi_1^1$, etc. can then be defined as usual over the class of computable predicates. We then say that an operator $\Gamma$ is computable (respectively $\Sigma_0^n$, arithmetical, etc.) if the relation $R_\Gamma$ is computable (respectively $\Sigma_0^n$, arithmetical, etc.). The following results are well-known.

**Theorem 2.1.** Let $\Gamma$ be an inductive operator.

(a) If $\Gamma$ is computable, then the sequence $\{\Gamma^n : n \in \omega\}$ is uniformly computable, $|\Gamma| \leq \omega$, and $Cl(\Gamma)$ is $\Sigma_0^1$.

(b) If $\Gamma$ is $\Sigma_0^n$, then $|\Gamma| \leq \omega$ and if $\Gamma$ is monotone $\Sigma_0^n$, then $Cl(\Gamma)$ is $\Sigma_0^n$.

(c) Any $\Sigma_0^n$ set is 1-1 reducible to the closure of some computable monotone operator.

(d) If $\Gamma$ is monotone arithmetical, then $|\Gamma| \leq \omega_1^{CK}$ (the least non-computable ordinal) and $Cl(\Gamma)$ is $\Pi_1^1$.

(e) Any $\Pi_1^1$ set is 1-1 reducible to the closure of a monotone $\Pi_0^1$ operator.

**Example 2.1.** The classic example of a computable monotone operator is given by the definition of the set of sentences of a propositional logic over an infinite set $a_0, a_1, \ldots$ of propositional variables. Identifying sentences $p, q$ with their Gödel number $gn(p), gn(q)$, we have for any $i, p, q, and A$:

(0) $a_i \in \Gamma(A)$,

(1) $\neg p \in \Gamma(A)$ if $p \in A$, and

(2) $p \land q \in \Gamma(A)$ if $p \in A$ and $q \in A$, and

(3) $p \in \Gamma(A)$ if $p \in A$.

Other clauses could be added to include disjunction, implication or other binary connectives. This operator is computable because for any sentence $p$, we can compute the (at most two) other sentences which need to be in $A$ for $p$ to get into $\Gamma(A)$. Similar computable inductive definitions can be given for the set of terms in a first order language and the set of formulas in predicate logic. In each case, the closure ordinal of such a $\Gamma$ is $\omega$ and the set of sentences (respectively, terms, formulas) is computable since for any sentence (term, formula) $p$ of length $n$, $p \in lfp(\Gamma) \iff p \in \Gamma^n$.

**Example 2.2.** Suppose we are given a computable or $\Sigma_0^1$ set $A_0$ of axioms for propositional logic together with the logical axioms $\neg p \lor p$ for each $p$ and a finite set of rules
as indicated below. Then the set of consequences of $A_0$ is generated by the operator $\Gamma$ where, for all sentences $p, q, r$ and all $A$:

(0) $p \in \Gamma(A)$ if $p$ is an axiom,
(1) $p \lor q \in \Gamma(A)$ if $p \in A$ or $q \in A$,
(2) $p \in \Gamma(A)$ if $p \lor p \in A$,
(3) $(p \lor q) \lor r \in \Gamma(A)$ if $p \lor (q \lor r) \in A$, and
(4) $q \lor r \in \Gamma(A)$ if $p \lor q \in A$ and $\neg p \lor r \in A$.

In this case, $\Gamma$ is a $\Sigma^0_1$ operator but is not computable since, for example, the Cut Rule (4) asks for the existence of a $p$ such that $p \lor q$ and $\neg p \lor r$ are in $A$.

Now in this particular case, the consequences of a computable set $A_0$ will be a computable set but a similar example can be given for first order logic where the consequences of a finite set of axioms for arithmetic is $\Sigma^0_1$ but not computable.

**Example 2.3.** The one-step provability operator for a computable Horn logic program is a $\Sigma^0_1$ monotone operator. That is, suppose $A$ is a computable set of propositional letters or atoms. We assume that $A = \mathbb{N}$. A logic programming clause is a construct of the form

$$C = p \leftarrow q_1, \ldots, q_m, \neg r_1, \ldots, \neg r_n$$

where $p, q_1, \ldots, q_m, r_1, \ldots, r_n$ are atoms. Given a clause $C$, we let

$$[C] = [p, \langle q_1, \ldots, q_m \rangle, \langle r_1, \ldots, r_n \rangle]$$

where by convention, we let $\langle q_1, \ldots, q_m \rangle = 0$ if $m = 0$ and $\langle r_1, \ldots, r_n \rangle = 0$ if $n = 0$. The atoms $q_1, \ldots, q_m, \neg r_1, \ldots, \neg r_n$ form the body of $C$ and the atom $p$ is its head. Given a set of atoms $M \subseteq A$, we say $M$ is a model of $C$ if either (i) there is an $q_i$ such that $q_i \notin M$ or there is an $r_j$ such that $r_j \in M$ ( $M$ does not satisfy the body of $C$) or (ii) $p \in M$ ($M$ satisfies the head of $C$). The clauses $C$ where $n = 0$ are called Horn clauses.

A program $P$ is a set of clauses. We say that $P$ is computable ($\Sigma^0_1$, arithmetical, etc.) if $\{[C] : C \in P\}$ is computable ($\Sigma^0_1$, arithmetical, etc.). A program entirely composed of Horn clauses is called a Horn program. If $P$ is a Horn program, then there is a one step provability operator associated with $P$, $T_P : \mathcal{P}(\mathbb{N}) \rightarrow \mathcal{P}(\mathbb{N})$, which is defined by

$$T_P(A) \text{ equals the set of all } p \text{ such that there exists a clause } C = p \leftarrow q_1, \ldots, q_n \text{ in } P \text{ such that } q_1, \ldots, q_n \in A.$$ 

A Horn program always has a least model which is the closure of $T_P$. It is the intended semantics of such a program.

For programs with bodies containing the negation operator $\not$, we will use the stable model semantics. Following (Gelfond and Lifschitz 1988), we define a stable model of the program as follows. Assume $M$ is a collection of atoms. The Gelfond-Lifschitz reduct of $P$ by $M$ is a Horn program arising from $P$ by first eliminating those clauses in $P$ which contain $\neg r$ with $r \in M$. In the remaining clauses, we drop all negative literals from the body. The resulting program $GL_M(P)$ is a Horn program. We call $M$ a stable
model of \( P \) if \( M \) is the least model of GL\(_M\)(\( P \)). In the case of a Horn program, there is a unique stable model, namely, the least model of \( P \). Alternatively, one can define a one step provability operator \( T_{P,M} \) relative to a logic program \( P \) consisting of clauses of the form of (1) and a collection of atoms \( M \) by defining \( T_{P,M}(A) \) to be the set all \( p \) such that there exists a clause \( C = p ← q_1, \ldots, q_n, ¬r_1, \ldots, ¬r_m \) in \( P \) such that (i) \( \{q_1, \ldots, q_n\} \subseteq A \) and (ii) \( \{r_1, \ldots, r_m\} \cap M = \emptyset \). Then \( M \) is a stable model if and only the closure of \( T_{P,M} \) equals \( M \). In general, if \( M \) is a computable set, then \( T_{P,M} \) is a monotone \( \Sigma_1^0 \) operator.

It should be pointed out that both Example 1 and Example 2 can reformulated in the framework of logic programming as computable Horn programs. That is, the set of rules is a computable set, even though the corresponding inductive operator need not be computable.

**Example 2.4.** Another setting where computable inductive operators arise is in computable algebra and computable model theory. Surveys on various topics in computable algebra and model theory can be found in (Ershov et. al 1998a; Ershov et. al 1998b).

A generic example of computable inductive operators that arise in computable algebra are effective closure systems which were introduced by Remmel (Remmel 1980). An effective closure system \( \mathcal{M} = (M, cl) \) consists of a computable set \( M \) of the natural numbers \( \mathbb{N} \) together with an operation \( cl : \mathcal{P}(M) \to \mathcal{P}(M) \), where \( \mathcal{P}(M) \) denotes the power set of \( M \), which satisfies the following:

(i) \( A \subseteq cl(A) \),

(ii) \( A \subseteq B \) implies \( cl(A) \subseteq cl(B) \),

(iii) \( cl(cl(A)) = cl(A) \), and

(iv) \( x \in cl(A) \) implies that for some finite \( A' \subseteq A \), \( x \in cl(A') \).

Furthermore we require that \( cl \) is effective on (indices of) finite sets. That is, we assume that there is an effective algorithm which, given \( x, y_1, \ldots, y_n \in M \), will decide whether or not \( x \in cl(y_1, \ldots, y_n) \), where \( cl(y_1, \ldots, y_n) \) denotes \( cl(\{y_1, \ldots, y_n\}) \). Note that this condition plus conditions (i)-(iv) ensure that such closure operators are at least \( \Sigma_1^0 \) monotone operators.

We also assume that \( (\mathcal{M}, cl) \) always satisfy the nontriviality axiom (v) below.

(v) \( cl(\emptyset) \neq \emptyset. \)

Here we write \( A =^* B \) if there exists a finite sets, \( E \) and \( F \), such that \( cl(A∪E) = cl(B∪F) \). Similarly we write that \( A \subseteq^* B \) if there is a finite set \( F \) such that \( A \subseteq cl(B∪F) \).

We say \( V \) is a substructure of \( \mathcal{M} \) or \( V \) is closed if \( V \subseteq M \) and \( cl(V) = V \). It is easy to see that both the set of c. e. substructures and the set of all substructures of \( \mathcal{M} \) form a lattice, where the meet operation is just the set theoretic intersection and the join of two substructures \( V \) and \( W \), denoted \( V + W \), is given by \( V + W = cl(V∪W) \). We let \( L(\mathcal{M}) \) denote the lattice of c. e. substructures of \( \mathcal{M} = (M, cl) \) and \( S(\mathcal{M}) \) the lattice of all substructures of \( \mathcal{M} \).

If \( \mathcal{M} \) also satisfies

(vi) (exchange) \( x \in cl(A∪\{y\}) – cl(A) \) implies \( y \in cl(A∪\{x\}) \),
we say \( \mathcal{M} \) is an effective Steinitz system. Effective Steinitz systems have been extensively studied, see (Nerode-Remmel 1982; Nerode-Remmel 1983), (Downey 1983a; Downey 1983b), and (Baldwin 1982; Baldwin 1984)).

Another natural class of examples are effective algebras. These are obtained as follows. Let \((M, R)\) be an effective universal algebra in the sense that \(M\) is a computable set and \(R\) a computable set of uniformly computable operations on \(M\). Then we naturally associate an effective closure system \((M, cl_R)\) with \((M, R)\) by setting \(cl_R(A)\) to be the closure of \(A\) under the operations of \(R\) and their projections. We call an effective closure systems \(M\) formed in this way an effective algebra. As we shall see most natural examples such as groups, rings, fields, vector spaces, etc. are effective algebras.

We remark that not all effective closure systems are effective algebras. For example, for any effective closure system \(M = (M, cl)\), we can define an intersection subsystem \((A, cl^*_A)\) for \(A \subseteq M\) where for any \(B \subseteq A\),

\[
cl^*_A(B) = cl(B) \cap A.
\]

It is easy to check that \((A, cl^*_A)\) is an effective closure system, but not necessarily an effective algebra.

We end this example with a partial list of some specific examples of effective closure systems that have been studied extensively in the literature. In particular, there has been considerable work on the lattice of c. e. substructures of various structures. Details can be found in the survey article by Nerode and Remmel (Nerode-Remmel 1985). Some general results on the lattice of substructures of effective closure systems can be found in the work of Downey and Remmel (Downey-Remmel 1998). Here we shall only give a brief description of the closure systems and we refer the reader to (Nerode-Remmel 1985) or (Downey-Remmel 1998) for more details.

**Sets.** Let \(M = (\omega, cl)\) where \(cl(A) = A\). In this case, \(L(M)\) is the lattice of c. e. sets. Clearly, \(cl\) is computable monotone operator in this case.

**Vector Spaces.** Let \(V_\infty\) denote a fully effective infinite dimensional vector space over a computable field. That is, \(V_\infty\) consists of a computable subset \(U\) of \(\omega\) and computable operations for addition and scalar multiplication on \(V_\infty\). Moreover we assume that \(V_\infty\) has an effective dependence algorithm, that is, there is a uniform algorithm which given any \(x, y_1, \ldots, y_n\) in \(U\), decides whether or not \(x \in (\{y_1, \ldots, y_n\})^*\) where \((A)^*\) denotes the subspace generated by \(A\). In this case, \(cl(A) = (A)^*\) and \(L(V_\infty)\) is the lattice of c. e. subspaces.

In this case, \(cl\) is a \(\Sigma_1^0\) monotone operator but it is not computable. That is, it is a result of Dekker (Dekker 1971) that every c. e. subspace \(V\) of \(V_\infty\) has a computable basis \(B\). Thus since there are c. e. subspaces which are not computable, it follows that the relation \(R_{cl}\) is only \(\Sigma_1^0\). Similar results hold for the remaining examples of closure operators given below.

**Fields.** Here \(F_\infty\) denotes a fully effective algebraically closed field with infinite computable transcendence base. Here \(cl(A)\) denote the algebraic closure of \(A\).

**Affine Spaces.** In this case \(M = (V_\infty, K\ell)\) where \(V_\infty\) a computable vector space over a computable ordered field. Define \(y \in K\ell(y_1, \ldots, y_n)\) if and only if \(y = \Sigma \lambda_i y_i\) with
Inductive Definitions

Σλ₁ = 1. Again this is a Steinitz algebra. We denote its lattice of c. e. affine subspaces by \( L(V_\infty, K\ell) \) to distinguish it from \( L(V_\infty) \) (cf. (Downey 1983b)).

**Locally Computable Rings and Modules.** There are many other computable rings and modules which are effective closure systems. For example, consider \( G = \bigoplus_{i \in \omega} \mathbb{Z} \), the free Abelian group on \( \omega \) generators.

**Subalgebras of Boolean Algebras.** (Remmel 1978; Remmel 1980) A computable Boolean algebra \( B = (B, \vee_B, \wedge_B, \neg_B) \) consists of a computable subset \( B \) of \( \omega \) and computable operations for the meet, \( \wedge_B \), join, \( \vee_B \), and complement, \( \neg_B \) operations which turn \( B \) into a Boolean algebra. In this case, \( \text{cl}(A) \) is the subalgebra generated by \( A \).

**Convex sets, \( K(V_\infty) \).** Finally, consider the structure \( K(V_\infty) = (V_\infty, \langle \rangle) \) from Kalantari (Kalantari 1981) and Downey (Downey 1984). Here we consider \( V_\infty \) where the underlying field is the rationals, \( \mathbb{Q} \), and \( \langle \rangle \) is the operation of taking the convex hull, viz,

\[
\{x_1, \ldots, x_n\} = \{y | y = \Sigma \lambda_i x_i \text{ with } \Sigma \lambda_i = 1 \text{ and } 0 \leq \lambda_i \leq 1\}.
\]

Then \((V_\infty, \langle \rangle)\) is obviously an effective closure system.

We note that in all the structures above, we can generate many classes of \( \Sigma^0_1 \) inductive operators by simply letting \( A \) be any computable or c. e. subset of the structure and defining a new closure operator \( \Gamma_A \) by defining \( \Gamma_A(S) = \text{cl}(A \cup S) \).

**3. Index sets for \( \Sigma^0_1 \) and computable monotone operators**

An important property of \( \Sigma^0_1 \) monotone operators \( \Gamma \) is that the relation \( m \in \Gamma(A) \) depends only on positive information about \( A \). That is, we have the following lemma, see (Hinman 1978), pg. 92.

**Lemma 3.1.** For any \( \Sigma^0_1 \) monotone operator \( \Gamma \), there is a computable relation \( R \) such that for all \( m \in \mathbb{N} \) and \( A \in \mathcal{P}(\mathbb{N}) \),

\[
m \in \Gamma(A) \iff (\exists n)(D_n \subseteq A \& R(m, n))
\]

(2)

It follows from Lemma 3.1 that the \( \Sigma^0_1 \) monotone inductive operators may be effectively enumerated as \( \Gamma_0, \Gamma_1, \ldots \) in the following manner. For all \( e, m \in \mathbb{N} \) and all \( A \in \mathcal{P}(\mathbb{N}) \), let

\[
m \in \Gamma_e(A) \iff (\exists n)[D_n \subseteq A \text{ and } \langle m, n \rangle \in W_e].
\]

**Lemma 3.2.**

(a) There is a primitive recursive function \( f \) such that for all \( m, e, a \):

\[
\Gamma_e(W_a) = W_{f(e, a)}.
\]

(b) The relation \( m \in \Gamma_e^i \) is \( \Sigma^0_i \) in \( m, e, t \).

(c) The relation \( m \in \text{lf}(\Gamma_e) \) is \( \Sigma^0_i \) in \( m, e \).

(d) There is a computable function \( h \) such that \( \text{lf}(\Gamma_e) = W_{h(e)} \).
Theorem 3.1. Fix an infinite c.e. set \( W \). Then \( \{ e : W \cap \text{lf}(\Gamma_e) \text{ is computable} \} \) is \( \Sigma^0_3 \) complete.

Proof. We make use of the well-known fact (Soare 1987) that \( \text{Rec} = \{ e : W_e \text{ is computable} \} \) is \( \Sigma^0_3 \) complete. Let \( \psi \) be a computable function such that \( W \cap W_e = W_{\psi(e)} \) for all \( e \). Now let \( C = \{ e : W \cap \text{lf}(\Gamma_e) \text{ is computable} \} \) and let \( h \) be the computable function defined in the proof of part \((d)\) of Lemma 3.2. Then \( e \in C \iff \psi(h(e)) \in \text{Rec} \), so that \( C \) is a \( \Sigma^0_3 \) set.

For completeness, first consider the case where \( W = \mathbb{N} \). We can use the \( s\)-\( m\)\,-\( n \) theorem to obtain a 1:1 computable function \( g \) that

\[
(m, s) \in \Gamma_{\psi(e)}(A) \iff m \in W_{e,s} \text{ or } (m, s + 1) \in A.
\]

It is easy to see that \( \text{lf}(\Gamma_{g(e)}) = W_e \times \mathbb{N} \) so that \( W_e \) is computable if and only if \( \text{lf}(\Gamma_{g(e)}) \) is computable. Hence \( g \) witnesses that \( \text{Rec} \) is 1:1 reducible to \( C \) since \( e \in \text{Rec} \iff g(e) \in C \). Thus \( C \) is \( \Sigma^0_3 \) complete.

For an arbitrary infinite c.e. set \( W \), let \( R \) be an infinite computable subset of \( W \) and let \( f \) be an increasing, computable function such that \( R = \{ f(0), f(1), \ldots \} \). Then for any \( e \), let \( W_{g(e)} = \{ f(i) : i \in W_e \} \) and observe that \( W_{g(e)} \subseteq W \) for all \( e \) and that \( W_{g(e)} \) is computable if and only if \( W_e \) is computable. It follows that \( W_e \) is computable if and only if \( W \cap \text{lf}(\Gamma_{g(e)}) \) is computable. Thus \( g \circ p \) shows that, in general, \( \text{Rec} \) is 1:1 reducible to \( C \) so that \( C \) is \( \Sigma^0_3 \)-complete for all \( W \).

Computable operators are continuous and we can use the indexing of (Cenzer and Remmel 1999), pg. 135, to define the \( e \)-th computable monotone operator \( \Delta_e \) for \( e \) in the \( \Pi^0_2 \) set of indices such that \( \phi_e \) is a total function. That is, let \( \sigma_0, \sigma_1, \ldots \) enumerate the set \( \{0, 1\}^* \) of finite strings of 0’s and 1’s. For \( \sigma, \tau \in \{0, 1\}^* \), we write \( \sigma \subseteq \tau \) if \( \sigma \) is an initial segment of \( \tau \) and we write \( \sigma \subseteq \tau \) if \( \{ i : \sigma(i) = 1 \} \subseteq \{ i : \tau(i) = 1 \} \). Then the partial computable function \( \phi_e : \mathbb{N} \to \mathbb{N} \) defines a computable monotone operator \( \Delta_e : \mathcal{P}(\mathbb{N}) \to \mathcal{P}(\mathbb{N}) \) if it satisfies the following four conditions.

1. \( (\forall m)(\exists n)[\phi_e(m) = n] \), that is, \( \phi_e \) is total.
2. \( (\forall m)(\forall n)[\sigma_m \subseteq \sigma_n \implies \sigma_{\phi_e(m)} \subseteq \sigma_{\phi_e(n)}] \).
Inductive Definitions

(3) \((\forall m)(\exists n)(\forall \sigma_i \in \{0,1\}^n)[|\sigma_{\phi_e(i)}| \geq m]\).

(4) \((\forall m)(\forall n)[\sigma_m \subseteq \sigma_n \implies \sigma_{\phi_e(m)} \subseteq \sigma_{\phi_e(n)}]\).

The first three clauses above simply define the set of indices of computably continuous functions from \(\{0,1\}^N \to \{0,1\}^N\). Then clause (4) ensures that the resulting operator is monotone. Let ICM denote the set of indices \(e\) which satisfy (1)–(4). For \(A \subseteq \mathbb{N}\) and \(n \in \mathbb{N}\), identify \(A\) with its characteristic function and let \(A_n = i\) where \(\sigma_i = A[n = (A(0), A(1), \ldots, A(n-1))].\) Then we may define the \(e\)-th computable monotone operator by declaring that

\[ m \in \Delta_e(A) \iff (\exists n)(\forall \sigma_i \in \{0,1\}^n)[|\sigma_{\phi_e(i)}| \geq m \& \sigma_{\phi_e(A_n)}(m) = 1]. \tag{3} \]

Note that if \(\phi_e\) satisfies conditions (1)–(4), then \(\Delta_e(A)\) also has a \(\Pi_0^0\) definition, namely,

\[ m \in \Delta_e(A) \iff (\forall n)(\forall \sigma_i \in \{0,1\}^n)[|\sigma_{\phi_e(i)}| \geq m] \implies \sigma_{\phi_e(A_n)}(m) = 1. \tag{4} \]

**Theorem 3.2.** The set ICM of indices of computable monotone operators is \(\Pi_2^0\) complete.

*Proof.* It is clear that ICM is a \(\Pi_2^0\) set. For the completeness, we define a reduction of the \(\Pi_2^0\) complete set \(Tot = \{e : \phi_e\ \text{is total}\}\) to ICM as follows. Let \(f\) be the computable function such that for any \(i, \phi_f(i)(i) = j\) where \(\sigma_j = (\phi_0(0), \phi_e(1), \ldots, \phi_e(|\sigma_i| - 1))\). Now if \(e \notin Tot\), then clearly \(\phi_f(e)\) is not total and, hence, \(f(e) \notin ICM\). However, if \(e \in Tot\), then it is easy to see that for all \(A, \Delta_f(e)(A) = \{m : \phi_e(m) = 1\}\) and, hence, \(\Delta_f(e)\) is a computable monotone operator. Thus \(e \in Tot \iff f(e) \in ICM.\)

*Lemma 3.3.* There is a primitive computable function \(g\) such that for all \(e \in ICM, \Delta_e = \Gamma_{g(e)}\).

*Proof.* Define \((m, n) \in W_{g(e)}\) if and only \((\exists k)(\forall \sigma_i \in \{0,1\}^k)[|\sigma_{\phi_e(i)}| > m]\) and there exists \(\sigma_i \in \{0,1\}^k\) such that \(\sigma_{\phi_e(i)}(m) = 1\) and \(\{j : \sigma_i(j) = 1\} \subseteq D_n\). We now verify that \(\Delta_e = \Gamma_{g(e)}\) if \(e \in ICM\).

Suppose first that \(m \in \Delta_e(A)\). Then find the least \(k\) such that \((\forall \sigma_i \in \{0,1\}^k)[|\sigma_{\phi_e(i)}| > m]\). Thus for \(\sigma_i = A[k], we have \sigma_{\phi_e(i)}(m) = 1. Now let \(D_n = A \cap \{0,1, \ldots, k-1\}\) \(\iff \{j < k : \sigma_i(j) = 1\}\).

It follows that \((m, n) \in W_{g(e)}\) so that \(m \in \Gamma_{g(e)}(A)\).

Next suppose that \(m \in \Gamma_{g(e)}(A)\) and let \(n, k\) and \(\sigma_i \in \{0,1\}^k\) be given as above so that \(\sigma_{\phi_e(i)}(m) = 1\) and \(\{j : \sigma(j) = 1\} \subseteq D_n \subseteq A\). It follows from clause (4) above that \(\sigma_{\phi_e(A_n)}(m) = 1\) and therefore \(m \in \Delta_e(A)\).

Thus we have shown that, for all \(A, \Delta_e(A) = \Gamma_{g(e)}(A)\) and hence \(\Delta_e = \Gamma_{g(e)}\).

*Lemma 3.4.*

(a) There is a partial computable function \(\delta\) such that for all \(m, e, a\) with \(a \in Tot\) and \(e \in ICM, \delta(e, a) \in Tot\) and \(\Delta_e(\{m : \phi_a(m) = 1\}) = \{m : \phi_{\delta(e,a)}(m) = 1\}\).

(b) There is a partial computable function \(\psi\) such that for all \(e, t\) with \(e \in ICM, \phi_{\psi(e,t)}\) is the characteristic function of \(\Delta_e^{+}\).
(c) There is a \( \Sigma^0_1 \) relation \( S \) such that

\[
m \in \text{lfp}(\Delta_e) \iff (m, e) \in S.
\]

**Proof.** (a) To compute \( \phi_{\delta(e,a)}(m) \), first find \( k \) so that \( |\sigma_{\delta(i)}| > m \) for all \( \sigma_i \in \{0, 1\}^k \). Then let \( \sigma = (\phi_{\alpha(0)}, \phi_{\alpha(1)}, \ldots, \phi_{\alpha(k - 1)}) \) and set \( \phi_{\delta(e,a)}(m) = \sigma_{\delta(i)}(m) \).

Parts (b) and (c) follow easily.

This shows that the closure of any computable monotone inductive operator is a c. e. set. In (Cenzer 1978), the first author considered the converse problem of whether any c. e. set is the closure of some computable monotone inductive operator. It is shown there that not every c. e. set is the closure of such an operator, but the every c. e. set is one-one reducible to such a closure. Here is an index set version of that result.

**Theorem 3.3.** There are primitive recursive functions \( f \) and \( g \) such that for all \( e \) and \( m, f(e) \in ICM \) and \( m \in W_e \iff g(m) \in \text{lfp}(\Delta_{f(e)}) \).

**Proof.** Define the computable monotone inductive operator \( \Delta_{f(e)} \) by

\[
(m, s) \in \Delta_{f(e)}(A) \iff [m \in W_{e,s} \lor (m, s + 1) \in A].
\]

It is easy to see that \( \text{lfp}(\Delta_{f(e)}) = \{ (m, s) : m \in W_e \} \), so that for any \( m \) and \( e \),

\[
m \in W_e \iff (m, 0) \in \text{lfp}(\Delta_{f(e)}).
\]

Thus we can take \( g(m) = (m, 0) \).

The index set complexity for \( \Sigma^0_1 \) operators given in Theorem 3.1 easily carries over for computable monotone operators since the operator \( \Gamma_{g(e)} \) defined in the proof is uniformly computable. Thus we have the following.

**Theorem 3.4.** \( \{ e : \text{lfp}(\Delta_e) \text{ is computable} \} \) is \( \Sigma^0_1 \) complete.

For the remainder of this section, we consider the complexity of two types of index sets associated with monotone operators. The first type comes from the cardinality of the least fixed point. For example, we will determine the complexity of the problem of whether \( \text{lfp}(\Gamma_e) \) is a finite or an infinite set. The second type comes from the closure ordinal of the operator. For example, we will determine the complexity of the problem of whether the closure ordinal of \( \Delta_e \) is finite or equals \( \omega \). For the remaining results in this section, we will omit the routine verifications of the complexity upper bounds.

**Theorem 3.5.** \( \{ e : |\Gamma_e| > 0 \} \) and \( \{ e : |\Gamma_e| = 0 \} \) is \( \Sigma^0_1 \) complete and \( \{ e : |\Gamma_e| = 0 \} \) is \( \Pi^1_0 \) complete.

**Proof.** For the completeness, let \( E \) be an arbitrary c. e. set and define a computable function \( f_E \) so that

\[
m \in \text{Gamma}(A) \iff (m = 0 \& e \in E).
\]

Clearly if \( e \notin E \), then \( |\Gamma_{f_E(e)}| = 0 \) and \( \text{lfp}(\Gamma_{f_E(e)}) = \emptyset \) and if \( e \in E \), then \( |\Gamma_{f_E(e)}| = 1 \) and \( \text{lfp}(\Gamma_{f_E(e)}) = \{0\} \). Thus \( f_E \) shows that the arbitrary \( \Sigma^0_1 \) set \( E \) is 1:1 reducible to \( \{ e : |\Gamma_e| > 0 \} \) and at the same time \( \mathbb{N} - E \) is 1:1 reducible to \( \{ e : |\Gamma_e| = 0 \} \).
A set is said to be d. c. e. if it is a difference of two c. e. sets.

**Theorem 3.6.** For any natural number \( k > 0 \),

(a) \( \{ e : \text{card}(lfp(\Gamma_e)) > k \} \) is \( \Sigma^0_1 \) complete and \( \{ e : \text{card}(lfp(\Gamma_e)) \leq k \} \) is \( \Pi^0_1 \) complete.

(b) \( \{ e : \text{card}(lfp(\Gamma_e)) = k \} \) is d. c. e. complete.

**Proof.** (a) For the completeness, modify the definition of \( f_E \) in the proof of Theorem 3.5 so that

\[
m \in \Gamma_{f_E(e)}(A) \iff [m \leq k \& e \in E].
\]

Then \( lfp(\Gamma_{f_E(e)}) = \emptyset \) if \( e \notin E \) and \( lfp(\Gamma_{f_E(e)}) = \{0, 1, \ldots, k\} \) if \( e \in E \). Again \( f_e \) shows that \( E \) is 1:1 reducible to \( \{ e : \text{card}(lfp(\Gamma_e)) > k \} \) and, hence, \( \{ e : \text{card}(lfp(\Gamma_e)) > k \} \) is \( \Sigma^0_1 \) complete.

(b) Clearly, \( \{ e : \text{card}(lfp(\Gamma_e)) = k \} \) is \( \{ e : \text{card}(lfp(\Gamma_e)) \leq k \} - \{ e : \text{card}(lfp(\Gamma_e)) < k \} \). For completeness, we need only show that for any c. e. sets \( C \) and \( D \) with \( D \subseteq C \), there is 1:1 computable function \( g \) such that \( e \in C - D \iff g(e) \in \{ e : \text{card}(lfp(\Gamma_e)) = k \} \). So let \( C \) and \( D \) be c. e. sets where \( D \subseteq C \) and define \( g \) so that

\[
m \in \Gamma_{g(e)}(A) \iff [(m < k \& e \in C) \lor (m = k \& k - 1 \in A \& e \in D)].
\]

If \( e \notin C \), then \( lfp(\Gamma_{g(e)}) = \emptyset \). If \( e \in C - D \), then \( |\Gamma_{g(e)}| = k \) and \( lfp(\Gamma_{g(e)}) = \{0, 1, \ldots, k - 1\} \). If \( e \in C \cap D \), then \( |\Gamma_{g(e)}| = 2 \) and \( lfp(\Gamma_{g(e)}) = \{0, 1, \ldots, k\} \).

**Theorem 3.7.**

(a) \( \{ e : lfp(\Gamma_e) \text{ is finite} \} \) is \( \Sigma^0_2 \) complete and \( \{ e : lfp(\Gamma_e) \text{ is infinite} \} \) is \( \Pi^0_3 \) complete.

(b) \( \{ e : lfp(\Gamma_e) \text{ is cofinite} \} \) is \( \Sigma^0_3 \) complete and \( \{ e : lfp(\Gamma_e) \text{ is coinfinite} \} \) is \( \Pi^0_3 \) complete.

**Proof.** This follows easily from the facts that \( \{ e : W_e \text{ is finite} \} \) is \( \Sigma^0_2 \) complete and \( \{ e : W_e \text{ is cofinite} \} \) is \( \Sigma^0_3 \) complete by letting \( \Gamma_{f_1(e)}(A) = W_e \) for all \( A \).

The corresponding result for computable monotone operators is a corollary.

**Theorem 3.8.**

(a) \( \{ e : lfp(\Delta_e) \text{ is infinite} \} \) is \( \Pi^0_3 \) complete.

(b) \( \{ e : lfp(\Delta_e) \text{ is cofinite} \} \) is \( \Sigma^0_3 \) complete and \( \{ e : lfp(\Delta_e) \text{ is coinfinite} \} \) is \( \Pi^0_3 \) complete.

Next we consider the closure ordinal of a monotone inductive operator.

**Theorem 3.9.** For any natural number \( t \geq 1 \):

(a) \( \{ e : |\Gamma_e| > t \} \) is \( \Sigma^0_2 \) complete and \( \{ e : |\Gamma_e| \leq t \} \) is \( \Pi^0_2 \) complete.

(b) \( \{ e : |\Gamma_e| = 1 \} \) is \( \Pi^0_2 \) complete.

(c) \( \{ e : |\Gamma_e| = t + 1 \} \) is \( D^0_2 \) complete.

**Proof.** We will use the fact that \( Fin = \{ e : W_e \text{ is finite} \} \) is a \( \Sigma^0_2 \) complete set. We can define a 1:1 computable function \( f \) so that

\[
m \in \Gamma_{f(e)}(A) \iff m = 0 \lor (\exists n \leq m)(n \in A) \lor (\exists n \geq m)(n \in W_e).
\]
If $W_e$ is infinite, then $\Gamma^1_{f(e)} = \mathbb{N}$ and $|\Gamma_e| = 1$. If $W_e$ is finite, let $M$ be the largest element of $W_e \cup \{0\}$. Then $\Gamma^1_{f(e)} = \{0, 1, \ldots, M\}$, $\Gamma^2_{f(e)} = \mathbb{N}$ and therefore $|\Gamma_{f(e)}| = 2$. Thus $e \in \text{Fin} \iff f(e) \in \{e : |\Gamma_e| > 1\}$ which establishes completeness for part (a) when $t = 1$ and the completeness of part (b).

For the completeness in part (a), fix $t \geq 1$ and define a $1:1$ computable function $g$ such that $m \in \Gamma^t_{g(e)}(A)$ if and only if

$$m = 0 \lor (m < t \land m - 1 \in A) \lor (m \geq t \land [(\exists n \geq m)(n \in W_e) \lor (t-1 \in A)].$$

Then it is easy to see that if $W_e$ is infinite, then for all $i$,

$$\Gamma^i_{g(e)} = \{x : x < i \lor x \geq t\},$$

so that $|\Gamma^i_{g(e)}| = t$ and $lfp(\Gamma^i_{g(e)}) = \mathbb{N}$. However if $W_e$ is finite and $M$ is the largest element of $W_e$, then for $i \leq t$,

$$\Gamma^i_{g(e)} = \{x : x < i \lor t \leq x \leq M\}$$

and

$$\Gamma^{t+1}_{g(e)} = \mathbb{N},$$

so that $|\Gamma^t_{g(e)}| = t+1$. Thus $e \in \text{Fin} \iff g(e) \in \{e : |\Gamma_e| > t\}$.

For the completeness in part (c) in the case where $t = 1$, it suffices to define a computable function $h$ such that $|\Gamma^{h(a,b)}| = 2$ if and only if $W_a$ is finite and $W_b$ is infinite. Let $Ev$ denote the set of even numbers and $Od$ denote the set of odd numbers. First define $h_{(a,b)}$ so that

$$2m \in \Gamma^{h_{(a,b)}}(A) \iff m = 0 \lor (\exists n \geq m)(n \in A \lor (\exists n \geq m)(n \in W_a)).$$

Then by our argument in case (a), $Ev \subseteq \Gamma^{h_{(a,b)}}(A)$ if $W_a$ is infinite. If $W_a$ is finite and $M$ is the greatest element of $W_a \cup \{0\}$, then $\Gamma^{h_{(a,b)}} \cap Ev = \{2x : x \leq M\}$ and $Ev \subseteq \Gamma^{h_{(a,b)}}(A)$. We then complete the definition of $h$ so that

$$2m + 1 \in \Gamma^{h_{(a,b)}}(A) \iff [m = 0 \lor (\exists n \geq m)(n \in W_a)]$$

$$\lor [0 \in A \land (\exists n \geq m)(n \in W_b)]$$

$$\lor [0 \in A \land m > 0 \land (2m - 1 \in A)].$$

Now if $W_a$ is infinite, then $0d \subseteq \Gamma^{h_{(a,b)}}(A)$ and, hence, $\Gamma^{h_{(a,b)}} = \mathbb{N}$ and $|\Gamma_{h(a,b)}| = 1$. Next suppose that $W_a$ is finite and $M$ is the greatest element of $W_a \cup \{0\}$. Then our definition of $h$ ensures that $\Gamma^{h_{(a,b)}} \cap 0d = \{2x + 1 : x \leq M\}$ since $0 \notin \Gamma^{h_{(a,b)}}$. Now if $W_b$ is infinite, then $Od \subseteq \Gamma^{h_{(a,b)}}(A)$ so that $\Gamma^{h_{(a,b)}} = \mathbb{N}$ and $|\Gamma_{h(a,b)}| = 2$. Finally, if $W_b$ is finite and $B$ is the largest element of $W_a \cup W_b \cup \{0\}$, then $\Gamma^{h_{(a,b)}} \cap Od = \{2x + 1 : x \leq B\}$ and $2B + 3 \in \Gamma^{h_{(a,b)}}(A)$ so that $|\Gamma_{h(a,b)}| \geq 3$. This shows that $\{e : |\Gamma_e| = 2\}$ is $D_3$ complete.

For the general case of part (c), fix $t > 1$ and define $h$ so that

$$2m \in \Gamma^{h_{(a,b)}}(A) \iff m = 0 \lor (m < t \land m - 1 \in A)$$

$$\lor (m \geq t \land [(\exists n \geq m)(n \in W_e \lor 2(t-1) \in A)].$$

Then we can argue as in case (a) that $Ev \subseteq \Gamma^{h_{(a,b)}}(A)$ if $W_a$ is infinite. On the other hand,
if $W_a$ is finite and $M$ is the largest element of $W_a \cup \{0\}$, then

$$\Gamma^t_{h(a,b)} \cap Ev = \{0, \ldots, 2(t-1)\} \cup \{2x : x \geq t\} \text{ and}$$

$$\Gamma^{t+1}_{h(a,b)} \cap Ev = Ev.$$  

We then complete the definition of $h$ so that

$$2m + 1 \in \Gamma_{h(a,b)}(A) \iff [m = 0 \lor (\exists n \geq m)(n \in W_a)]$$

$$\lor [2(t - 1) \in A \land (\exists n \geq m)(n \in W_a)]$$

$$\lor [m > 0 \land 2(t - 1) \in A \land 2m - 1 \in A].$$

It can be verified that $W_a$ is finite and $W_b$ is infinite if and only if $\Gamma_{h(a,b)} = t + 1$.

**Theorem 3.10.** \{e : |\Gamma_e| = \omega\} is $\Pi^0_3$ complete and \{e : |\Gamma_e| < \omega\} is $\Sigma^0_3$ complete.

**Proof.** We use the $\Sigma^0_3$ completeness of $Cof = \{e : W_e \text{ cofinite}\}$. We define a 1:1 computable function $f$ so that $W_e$ is cofinite if and only if $|\Gamma_{f(e)}| < \omega$. Define $f$ so that

$$2n \in \Gamma_{f(e)}(A) \iff n = 0 \lor 2n - 2 \in A \lor 2n + 1 \in A;$$

$$2n + 1 \in \Gamma_{f(e)}(A) \iff (\exists m > n)(2m + 1 \in A)$$

$$\lor (\exists m < n)[2m \in A \land (\forall i \leq n)(m \leq i \longrightarrow i \in W_e)].$$

We make the following observations. First, $\Gamma^0_{f(e)} = \{0\}$ for all $e$. Next it is easy to see by the fact of our two conditions defining $f$ that we certainly have $2n \in \Gamma^m_{f(e)}$ for all $n$ and $e$ and, moreover, $2n \in \Gamma^t_{f(e)}$ for $n \geq t$ if and only if $2n + 1 \in \Gamma^t_{f(e)}$. Thus if $Ev$ is the set of even numbers, then $Ev \subseteq lfp(\Gamma_{f(e)})$ for all $e$.

Now fix $e$ and let $\Gamma = \Gamma_{f(e)}$. First suppose that $W_e$ is cofinite and $M$ is the least natural number such that $i \in W_e$ for all $i \geq M$. It follows from our second condition defining $f$ that, since $2M \in \Gamma^{M+1}$, $2n + 1 \in \Gamma^{M+2}$ for all $n \geq M$. But then it is easy to see that $2n + 1 \in \Gamma^{M+3}$ for all $n$ and $2n \in \Gamma^{M+4}$ for all $n$. Thus $lfp(\Gamma) = \mathbb{N}$ and $|\Gamma| \leq M + 4$. On the other hand, suppose that $|\Gamma| = k$ is finite. It follows that $2n \in \Gamma^k$ for all $n$. Let $t \leq k$ be the least such that $\{n : 2n \in \Gamma^t\}$ is infinite. By our observations above $t > 1$ so let $M$ be the maximum of $\{m : 2m \in \Gamma^{t-1}\}$. Thus for infinitely many $n \geq t$, $2n \in \Gamma^t$ and, hence, $2n + 1 \in \Gamma^{t-1}$. Now let $s$ be the least $k \leq t - 1$ such that $\{n : 2n + 1 \in \Gamma^k\}$ is infinite. Again it must be the case that $s > 1$ so that $\Gamma^{s-1}$ must be finite. Now let $p$ be the largest element such that $2p \in \Gamma^{s-1}$. Because $\{n : 2n + 1 \in \Gamma^s\}$ is infinite, it must be that case that for arbitrarily large $n$, there is an $m \leq p$ such that $2m \in \Gamma^{s-1}$ and $i \in W_e$ for $m \leq i \leq n$. But this implies that $W_e$ is cofinite.

The operator $\Gamma_{f(e)}$ defined in the proof Theorem 3.10 does not define a computable monotone operator so that we cannot conclude that $\{e : |\Delta_e| = \omega\}$ is $\Pi^0_3$ complete. In fact, $\{e : |\Delta_e| = \omega\}$ is $\Pi^0_3$ complete as our next result will show.

**Theorem 3.11.** \{e : e \in ICM \& |\Delta_e| = \omega\} is $\Pi^0_3$ complete.

**Proof.** We will define a 1:1 computable function $f$ such that for all $e$, $f(e) \in ICM$ and $W_e$ is finite if and only if $|\Delta_{f(e)}| < \omega$. The desired $f$ is the function defined in the
proof of Theorem 3.3 where
\[ \langle m, s \rangle \in \Delta_f(e)(A) \iff [m \in W_{a,s} \lor \langle m, s+1 \rangle \in A]. \]
Suppose first that \( W_e \) is infinite. Then there are arbitrarily large \( m \) and \( s \) such that \( m \in W_{e,s+1} - W_{e,s} \) and therefore \( \langle m, 0 \rangle \in \Delta_{f(e)}^{s+2} - \Delta_{f(e)}^{s+1} \). Thus \( |\Delta_{f(e)}| = \omega \). On the other hand, if \( W_e \) is finite, then there is a finite \( s \) such that \( m \in W_e \) implies \( m \in W_{e,s} \) for all \( m \). It follows that \( |\Delta_f(e)| \leq s + 1 \) and is finite.

4. Weakly finitary monotone operators

It follows from Lemma 2.1 that any \( \Sigma^0_1 \) monotone inductive operator \( \Gamma \) is finitary, that is, for any \( x \) and any set \( A \), we have \( x \in \Gamma(A) \) if and only if there is a finite subset \( D \) of \( A \) such that \( x \in \Gamma(D) \). The idea of a weakly finitary operator is to have a finite set \( m_1, \ldots, m_k \) of exceptional numbers which may be put into \( \Gamma(A) \) when an infinite set is included in \( A \). If there are exactly \( k \) exceptional numbers, then the operator \( \Gamma \) will be called \( k \)-weakly finitary. For example, we might allow some finite number of consequences of the \( \omega \)-rule in a subsystem of Peano arithmetic and still obtain a c. e. theory.

**Definition 4.1.**

1. We say that a monotone inductive operator \( \Gamma : \mathcal{P}(\mathbb{N}) \to \mathcal{P}(\mathbb{N}) \) is weakly finitary if there is a finite set \( S_\Gamma \) such that for all \( A, \)
   
   (a) \( x \notin S_\Gamma \) and \( x \in \Gamma(A) \) implies there exists a finite set \( F \subseteq A \) such that \( x \in \Gamma(F) \)
   and
   
   (b) \( x \in S_\Gamma \) and there is a family \( \mathcal{F}_{\Gamma,x} \) of subsets of \( \mathbb{N} \) which includes at least one infinite subset of \( \mathbb{N} \) such that \( x \in \Gamma(A) \) implies there exists an \( F \subseteq A \) such that \( x \in \Gamma(F) \) for some \( F \in \mathcal{F}_{\Gamma,x} \).

2. We say \( \Gamma = \Lambda_{k,e} \) is a \( k \)-weakly finitary \( \Sigma^0_1 \) monotone inductive operator with index \( \langle k, e \rangle = \langle k, \langle d, \langle m_1, c_1, \ldots, m_k, c_k \rangle \rangle \) if
   
   (i) \( \Gamma \) is a weakly finitary monotone operator with \( S_\Gamma = \{ m_1 < \cdots < m_k \} \),
   
   (ii) for all \( m_i \in S_\Gamma \), \( \mathcal{F}_{\Gamma,m_i} = \{ W_a : a \in W_e \} \),
   
   (iii) for all \( A \in \mathcal{P}(\mathbb{N}) \) and \( m \in \mathbb{N} \), \( m \in \Lambda_{k,e}(A) \) if and only if either
   
   (a) \( m \in \Gamma_d(A) \) or
   
   (b) for some \( i, m = m_i \) and \( (\exists a \in W_e)(W_a \subseteq A) \).

**Example 4.1.**

One example of this type of operator comes from attempts to extend logic programming to be able to reason about infinite sets described in (Cenzer, Marek and Remmel 2005). They defined an extension of logic programming which they call extended set-based programming (ESB). In this example, we shall give the formal definitions of ESB constraints, clauses, programs and define the analogue of Horn programs and stable models for ESB programs. The basic idea is to incorporate constraints involving infinite sets into logic programming clauses by using various types of indexing schemes.
Inductive Definitions

To describe the constraints used by Cenzer, Marek and Remmel, we first need to describe three types of indices for subsets of the natural numbers.

1. **Explicit indices of finite sets.** Recall that $D_n = \{x_1 < \cdots < x_k\}$ where $n = \sum_{i=1}^k 2^i$.

2. **Computable indices of computable sets.** Let $\phi_0, \phi_1, \ldots$ be an effective list of all partial computable functions. By a computable index of a computable set $R$, we mean an $e$ such that $\phi_e$ is the characteristic function of $R$. If $\phi_e$ is a total $\{0,1\}$-valued function, then $R_e$ will denote the set $\{x \in \mathbb{N} : \phi_e(x) = 1\}$.

3. **C. e. indices of c. e. sets.** By a c. e. index of a c. e. set $W$, we mean an $e$ such that $W$ equals the domain of $\phi_e$, that is, $W_e = \{x \in \mathbb{N} : \phi_e(x)\}$ converges.

No matter what type of indices we use, we shall always consider two types of constraints based on $X$ and a finite set of indices $F$, namely, $\langle X, F \rangle^= \subseteq \langle X, F \rangle^= \subseteq \langle X, F \rangle^\subseteq$. For any subset $\langle X, F \rangle^\subseteq$, we mean a model of $\langle X, F \rangle^\subseteq$, written $M \models \langle X, F \rangle^\subseteq$, if there exists $e \in \mathcal{F}$ such that $M \cap X$ equals the set with index $e$. Similarly, we shall say that $M$ is a model of $\langle X, F \rangle^\subseteq$, written $M \models \langle X, F \rangle^\subseteq$, if there exists an $e \in \mathcal{F}$ such that $M \cap X$ contains the set with index $e$.

Then Cenzer, Marek and Remmel consider three different types of constraints.

(A) **Finite constraints.** Here we assume that we are given an explicit index $x$ of a finite set $X$ and a finite family $\mathcal{F}$ of explicit indices of finite subsets of $X$. We shall identify the finite constraints $\langle X, \mathcal{F} \rangle^= \subseteq \langle X, \mathcal{F} \rangle^\subseteq$ with their codes, $\langle 0, 0, x, n \rangle$ and $\langle 0, 1, x, n \rangle$ respectively where $\mathcal{F} = D_n$. Here the first coordinate 0 tells that the constraint is finite, the second coordinate is 0 or 1 depending on whether the constraint is $\langle X, \mathcal{F} \rangle^\subseteq$ or $\langle X, \mathcal{F} \rangle^\subseteq$, and the third and fourth coordinates are the codes of $X$ and $\mathcal{F}$ respectively.

(B) **Computable constraints.** Here we assume that we are given a computable index $x$ of a computable set $X$ and a finite family $\mathcal{R}$ of computable indices of computable subsets of $X$. Again we shall identify the computable constraints $\langle X, \mathcal{R} \rangle^\subseteq \subseteq \langle X, \mathcal{R} \rangle^\subseteq$ with their codes, $\langle 1, 0, x, n \rangle$ and $\langle 1, 1, x, n \rangle$ respectively, where $\mathcal{R} = D_n$. Here the first coordinate 1 tells that the constraint is computable, the second coordinate is 0 or 1 depending on whether the constraint is $\langle X, \mathcal{R} \rangle^\subseteq$ or $\langle X, \mathcal{R} \rangle^\subseteq$, and the third and fourth coordinates are the codes of $X$ and $\mathcal{R}$ respectively.

(C) **C. e. constraints.** Here we are given a c. e. index $x$ of a c. e. set $X$ and a finite family $\mathcal{W}$ of c. e. indices of c. e. subsets of $X$. Again we identify the finite constraints $\langle X, \mathcal{W} \rangle^\subseteq$ with their codes, $\langle 2, 0, x, n \rangle$ and $\langle 2, 1, x, n \rangle$ respectively, where $\mathcal{W} = D_n$. The first coordinate 2 tells that the constraint is c. e., the second coordinate is 0 or 1 depending on whether the constraint is $\langle X, \mathcal{W} \rangle^\subseteq$ or $\langle X, \mathcal{W} \rangle^\subseteq$, and the third and fourth coordinates are the codes of $X$ and $\mathcal{W}$.

An extended set-based clause is defined to be a clause of the form

$$\langle X, \mathcal{A} \rangle^+ \leftarrow (Y_1, B_1)^\subseteq, \ldots, (Y_k, B_k)^\subseteq, (Z_1, C_1)^=, \ldots, (Z_l, C_l)^=, \tag{5}$$

where $*$ is either $\doteq$ or $\subseteq$. We shall refer to $\langle X, \mathcal{A} \rangle^+$ as the head of $C$, written $\text{head}(C)$, and $(Y_1, B_1)^\subseteq, \ldots, (Y_k, B_k)^\subseteq, (Z_1, C_1)^=, \ldots, (Z_l, C_l)^=$ as the body of $C$, written $\text{body}(C)$. 

Here either $k$ or $l$ may be 0. $M$ is said to be a model of $C$ if either $M$ does not model every constraint in $\text{body}(C)$ or $M \models \text{head}(C)$.

Again we shall talk about three different types of clauses.

(a) **Finite clauses.** These are clauses in which all of the constraints are finite constraints.

(b) **Computable clauses.** These are clauses where all the constraints appearing in the clause are finite or computable constraints and at least one constraint is a computable constraint.

(c) **C. e. clauses:** These are clauses where all the constraints appearing in the clause are finite, computable or c. e. constraints and there is at least one c. e. constraint.

An extended set-based (ESB) program $P$ is a set of clauses of the form of (1). We say that an ESB program $P$ is computable, if the set of codes of the clauses of $P$ is a computable set. Here the code of a clause $C$ of the form of (1) is $\langle c_1, e_1, \ldots, c_k, f_1, \ldots, f_l \rangle$ where $c$ is the code of $\langle X, A \rangle$, $e_1$ is the code of $\langle Y_{i_1}, B_{i_1} \rangle$ for $i = 1, \ldots, k$ and $f_j$ is the code of $\langle Z_j, C_j \rangle$ for $j = 1, \ldots, l$.

Given a program $P$, we let $\text{Fin}(P), (\text{Comp}(P), CE(P))$ denote the set of all finite (computable, c. e.) clauses in $P$. It is easy to see from our coding of clauses that if $P$ is a computable ESB program, then $\text{Fin}(P), \text{Comp}(P)$ and $CE(P)$ are also computable ESB programs.

Let $P$ be a computable ESB program. We will say that $P$ is **computable with finite constraints** if $P = \text{Fin}(P)$. Similarly we say that $P$ is **computable with computable constraints** if $P = \text{Fin}(P) \cup \text{Comp}(P)$ and $\text{Comp}(P) \neq \emptyset$, and $P$ is **computable with c. e. constraints** if $CE(P) \neq \emptyset$. Finally we say that $P$ is **weakly finite with computable constraints** if $P$ is computable with computable constraints and the set of heads of clauses in $\text{Comp}(P)$ is finite, and $P$ is **weakly finite with c. e. constraints** if $P$ is computable with c. e. constraints and the set of heads of clauses in $\text{Comp}(P) \cup CE(P)$ is finite.

Next we define the analogue of Horn programs for ESB programs. A Horn program $P$ is a set of clauses of the form

$$\langle X, A \rangle \subseteq \langle Y_1, B_1 \rangle \subseteq \ldots \subseteq \langle Y_k, B_k \rangle \subseteq \ldots$$

where $A$ is a singleton. We define the **one-step provability operator**, $T_P : 2^\mathbb{N} \to 2^\mathbb{N}$ so that for any $S \subseteq \mathbb{N}$, $T_P(S)$ is the union of the set of all $F_e$ such that there exists a clause $C \in P$ such that $S \models \text{body}(C), \text{head}(C) = \langle X, A \rangle$ and $A = \{e\}$ where $F_e = D_e$ if $\text{head}(C)$ is a finite constraint, $F_e = R_e$ if $\text{head}(C)$ is a computable constraint, and $F_e$ is $W_e$ if $\text{head}(C)$ is an c. e. constraint. It is easy to see that $T_P$ is a monotone operator and hence there is a least fixed point which we denote by $M^P$. Moreover it is easy to check that $M^P$ is a model of $P$.

If $P$ is an ESB Horn program in which the body of every clause consists of finite constraints, then one can easily that the least fixed point of $T_P$ is reached in $\omega$-steps, that is, $M^P = T_P \uparrow^\omega (\emptyset)$. However, if we allow clauses whose bodies contain either computable or c. e. constraints, then we can no longer guarantee that we reach the least fixed point of $T_P$ in $\omega$ steps. Here is an example.
Example 4.2. Let $e_n$ be the explicit index of the set $\{n\}$ for all $n \geq 0$, let $w$ be a computable index of $\mathbb{N}$ and $f$ be a computable index of the set of even numbers $E$. Consider the following program.

\[
\langle \{0\}, \{e_0\} \rangle \subseteq \leftarrow \langle \{2x + 2\}, \{e_{2x+2}\} \rangle \subseteq \langle \{2x\}, \{e_{2x}\} \rangle \subseteq \langle \{E\}, \{f\} \rangle
\]

Clearly $\mathbb{N}$ is the least model of $P$ but it takes $\omega + 1$ steps to reach the fixed point. That is, it is easy to check that $T_P \uparrow^\omega = E$ and that $T_P \uparrow^\omega+1 = \mathbb{N}$

Several results about ESB and weakly ESB programs were proved in (Cenzer, Marek and Remmel 2005). Their basic result about ESB Horn programs is the following.

Theorem 4.1.

(a) If $P$ is a computable ESB Horn Program with finite constraints, then the least fixed point of the one step provability operator $T_P$ is c. e..

(b) If $P$ is a weakly finite ESB Horn program with computable constraints such that $\text{Fin}(P)$ is computable, then the least fixed point of the one step provability operator $T_P$ is c. e..

(c) If $P$ is a weakly finite ESB Horn program with c. e. constraints such that $\text{Fin}(P)$ is computable, then the least fixed point of the one step provability operator $T_P$ is c. e..

In fact, a similar result to Theorem 4.1 holds for $k$-weakly $\Sigma_0^1$ monotone operators.

Theorem 4.2. Let $\Lambda$ be a $k$-weakly $\Sigma_0^1$ monotone operator with index \((k,e) = (k,d,\langle m_1, e_1, \ldots, m_k, e_k \rangle)\). Then

(a) $|\Lambda| \leq \omega \cdot (k + 1)$.

(b) $\text{lfp}(\Lambda)$ is $\Sigma_0^1$.

Proof. We will present an informal procedure which constructs the closure in $\leq k + 1$ rounds where each round may consist of as many $\omega$ steps.

Round 1. First let $U_0 = \text{lfp}(\Gamma_d)$. Since $\Gamma_d$ is a $\Sigma_0^1$ monotone inductive operator, $U_0$ is c. e. by Theorem 2.1. Next consider the finite set

\[
F_0 = \{m_i : (\exists a \in W_{e_i})(W_a \subseteq U_0)\}.
\]

We can not necessarily find $F_0$ effectively, but, nevertheless, $F_0$ is a finite set so that $A_1 = U_0 \cup F_0$ will be a c. e. set. If $F_0 = \emptyset$, then $\text{lfp}(\Lambda) = U_0$ and $|\Lambda| \leq \omega$. Otherwise go on to Round 2.

We now present the description of Round $n + 1$, for $n \geq 1$. Assume that $A_n$ is the result of step $n$.

Round $(n+1)$. Consider the set $U_n = \Gamma_d^\omega(A_n)$. It is easy to see that that since $A_n$ is c. e., $U_n$ is also c. e.. Next consider the finite set

\[
F_n = \{m_i : (\exists a \in W_{e_i})(W_a \subseteq U_n)\}.
\]

Again we can not necessarily find $F_n$ effectively, but, nevertheless, $A_{n+1} = U_n \cup F_n$ is a
It is easy to see that $\Lambda$ is a $k$-a the elements.

Finally, we complete the definition of $\Lambda$ by adding the following rules which govern when $\Lambda$ holds for all $i$.

Let $\Lambda$ be a $\Lambda$. If $\Lambda(\cdot)$.

Then $\Lambda(\cdot)$ must be completed after at most $k$ rounds, so that $|\Lambda| \leq \omega \cdot (k + 1)$ and $lfp(\Lambda)$ is always a c.e. set.

**Example 4.3.** It is easy to construct an example $\Lambda$ of a $k$-weakly finitary $\Sigma^0_1$ monotone operator with $|\Lambda| = \omega \cdot (k + 1)$. That is, let $A_0, \ldots, A_k$ be a set of infinite computable sets that partition $\mathbb{N}$. Let $A_i = \{a_{0,i} < a_{1,i} < \ldots\}$ for $i = 0, \ldots, k$. First define a $\Sigma^0_1$ monotone operator $\Gamma$ such that for all $A \subseteq \mathbb{N}$,

(i) $a_{0,0} \in \Gamma(A)$,

(ii) for all $j \geq 0$, $a_{j+1,0} \in \Gamma(A)$ if and only if $a_{j,0} \in A$,

(iii) for all $i \geq 1$, $a_{i,i} \in \Gamma(A)$ if and only if $a_{0,i} \in A$, and

(iv) for all $i \geq 1$ and $j \geq 1$, $a_{j+1,i} \in \Gamma(A)$ if and only if $a_{j,i} \in A$.

Finally, we complete the definition of $\Lambda$ by adding the following rules which govern when the elements $a_{0,1}, \ldots, a_{0,k}$ can be in $\Lambda(A)$.

For all $i > 0$, $a_{0,i} \in \Lambda(A)$ if and only if $A_{i-1} \subseteq A$.

It is easy to see that $\Lambda$ is a $k$-weakly finitary $\Sigma^0_1$ monotone operator and that

$$
\Lambda^\omega = A_0,
\Lambda^{\omega+1} = A_0 \cup \{a_{0,1}\},
\Lambda^{2\omega} = A_0 \cup A_1,
\Lambda^{2\omega+1} = A_0 \cup A_1 \cup \{a_{0,2}\},
\vdots
\Lambda^{k\omega} = A_0 \cup A_1 \cup \cdots \cup A_{k-1},
\Lambda^{k\omega+1} = A_0 \cup A_1 \cup \cdots \cup A_{k-1} \cup \{a_{0,k}\},
\Lambda^{(k+1)\omega} = A_0 \cup A_1 \cup \cdots \cup A_k = \mathbb{N}.
$$

Thus $|\Lambda| = \omega(k+1)$.

The following lemma gives an alternate approach to proving part (b) of Theorem 4.2 and will be needed below.

**Lemma 4.1.** Let $\Lambda$ be a $k$-weakly finitary $\Sigma^0_1$ monotone operator with index

$$
\langle k, e \rangle = \langle d, \langle m_1, e_1, \ldots, m_k, e_k \rangle \rangle.
$$

Then

(a) for some finite subset $F$ of $\{m_1, \ldots, m_k\}$, $lfp(\Lambda) = \Gamma^\omega_d(F)$ and

(b) for some finite subset $G$ of $\{m_1, \ldots, m_k\}$, $\Lambda^\omega = \Gamma^\omega_d(G)$.

**Proof.** For part (a), let $F = \{m_i : m_i \in lfp(\Lambda)\}$. Then certainly $\Gamma^\omega_d(F) \subseteq \Lambda^\omega(F) \subseteq lfp(\Lambda)$. For the reverse inclusion, it suffices to show that $C = \Gamma^\omega_d(F)$ is closed under $\Lambda$. If $\Lambda(C) = C \neq \emptyset$, then either (i) there is some $y \notin S_F = \{m_1, \ldots, m_k\}$ such that $y \in \Gamma^\omega_d(C) - C$ or (ii) there is some $m_i \notin C$ such $W_a \subseteq C$ for some $a \in W$. Note that (i) is not possible. That is, $\Gamma^\omega_d(C) \subseteq C$ because $\Gamma^\omega_d$ is a $\Sigma^0_1$ monotone operator and, hence,
Inductive Definitions

\[ \Gamma_d(\Gamma^\omega(F)) = \Gamma^\omega(F) \text{.} \]

But (ii) is not possible since otherwise \( m_i \in F \) and \( F \subseteq C \). Thus it must be the case that \( \Lambda(C) = \Gamma_d(C) \).

For part (b), let \( G = \{ m_i : m_i \in \Lambda^\omega \} \). Since \( G \) is a finite set, there is some finite \( t \) such that \( G \subseteq \Lambda^t \). Then certainly \( \Gamma^\omega_d(G) \subseteq \Lambda^\omega(G) \subseteq \Lambda^\omega(\Lambda^t) = \Lambda^\omega \). For the reverse inclusion, suppose \( D = \Gamma^\omega_d(G) \) and \( \Lambda^\omega - D \neq \emptyset \). Then let \( s \) be the least stage such that there is an \( x \in \Lambda^s - D \). Then either (I) \( x \notin S_t = \{ m_1, \ldots, m_k \} \) and hence, there is some finite set \( F \subseteq \Lambda^{t-1} \) such that \( x \in \Gamma(F) \) or (II) \( x = m_i \notin G \) and \( W_a \subseteq \Lambda^{t-1} \) for some \( a \in W_e \). Note that in case (I), \( F \subseteq D \) by our choice of \( s \). But since \( F \) is finite, there must be some finite \( t \) such \( F \subseteq \Gamma^t(G) \) so that \( x \in \Gamma(F) \subseteq \Gamma^t(G) \subseteq \Gamma^\omega(G) = D \). Thus case (I) cannot hold. But Case (II) is not possible since otherwise \( m_i \in G \) and \( G \subseteq D \). Thus it must be the case that \( \Lambda^\omega = \Gamma^\omega_d(G) \).

It is possible to develop a theory of index sets for weakly finitary \( \Sigma^0_\infty \) inductive operators. In general, this theory is more subtle than the corresponding theory of \( \Sigma^0_1 \) inductive operators. We will not attempt in this paper to prove analogues of all the index set results in Section 3. Instead, we will give a couple of examples of index set results for weakly finitary \( \Sigma^0_1 \) inductive operators where there is a contrast between the index set result for weakly finitary \( \Sigma^0_1 \) inductive operators and the corresponding index set result for \( \Sigma^0_1 \) inductive operators.

Clearly, \( \{ e : |\Gamma_e| \leq \omega \} = \mathbb{N} \) and is hence computable since for any \( \Sigma^0_1 \) inductive operator \( \Gamma, \Gamma^\omega = \text{lfp}(\Gamma) \). In contrast, for weakly finitary \( \Sigma^0_1 \) inductive operators we have the following.

**Theorem 4.3.**

(a) For all \( k \geq 1 \), the set of \( e \) such that \( \langle k, e \rangle = \langle k, \langle d, \langle m_1, e_1, \ldots, m_k, e_k \rangle \rangle \rangle \rangle \) and \( \{ m_1, \ldots, m_k \} \cap cl(\Lambda_{k,e}) = \emptyset \) (in which case \( cl(\Lambda_{k,e}) = \Gamma^\omega_d \)) is a complete \( \Pi^0_3 \) set.

(b) For all \( k \geq 1 \), \( \{ e : |\Lambda_{k,e}| \leq \omega & \{ m_1, \ldots, m_k \} \subseteq \Lambda_{k,e} \} \) is \( \Sigma^0_1 \) complete.

(c) For all \( k \geq 2 \), \( \{ e : |\Lambda_{k,e}| \leq \omega \} \) is \( D^0_0 \) complete.

**Proof.** For the upper bound for part (a), suppose \( \langle k, e \rangle = \langle k, \langle d, \langle m_1, e_1, \ldots, m_k, e_k \rangle \rangle \rangle \rangle \). Then it is easy to see from our construction in Theorem 4.2, that \( \{ m_1, \ldots, m_k \} \cap cl(\Lambda_{k,e}) = \emptyset \) only if there is no \( i \) and \( a \in W_e \) such that \( W_a \subseteq \Gamma^\omega_d \). Since \( \Gamma_d \) is a \( \Sigma^0_1 \) inductive operator, \( \Gamma^\omega_d \) is a c.e. set. Thus \( \{ m_1, \ldots, m_k \} \cap cl(\Lambda_{k,e}) = \emptyset \) if and only

\[
(\forall i \in \{ 1, \ldots, k \})(\exists a \in W_e)(\exists c) (e \in W_a & c \notin \Gamma^\omega_d)
\]

which is a \( \Pi^0_3 \) predicate.

Next we consider the upper bounds for parts (b) and (c). Fix a set \( F \subseteq \{ 1, \ldots, k \} \). For each index \( \langle k, e \rangle = \langle k, \langle d, \langle m_1, e_1, \ldots, m_k, e_k \rangle \rangle \rangle \rangle \), let \( M_{F,k,e} = \Gamma^\omega_d(\{ m_i : i \in F \}) \).

Now fix \( \langle k, e \rangle \). By Lemma 4.1, we know there is some \( F \) such that \( M_{F,k,e} = \Lambda_{k,e}^\omega \).

We are interested in analyzing the predicate that

\[
Q(F,k,e) : M_{F,k,e} = \Lambda_{k,e}^\omega
\]

First suppose that \( F,G \subseteq \{ 1, \ldots, k \} \) and \( M_{F,k,e}, M_{G,k,e} \subseteq \Lambda_{k,e}^\omega \). Then it is easy to see...
that there must be some finite stage \( t \) such that \( G \cup F \subseteq \Lambda^t_{k,e} \). But then
\[
M_{F \cup G,k,e} = \Gamma_d(G \cup F) \\
\subseteq \Gamma_d(\Lambda^t_{k,e}) \\
\subseteq \Lambda^t_{k,e}(\Lambda^t_{k,e}) \\
= \Lambda^t_{k,e}.
\]

It thus follows that a particular \( F \) such that \( M_{F,k,e} = \Lambda^t_{k,e} \) is the maximal \( G \) such that \( M_{G,k,e} \subseteq \Lambda^t_{k,e} \).

Now if \( M_{F,k,e} = \Lambda^t_{k,e} \), then we can list the elements of \( F \) in the order in which they appear in the sequence \( \{\Lambda^i_{k,e}\}_{i \geq 0} \). That is, there is listing of \( F = \{f_1, \ldots, f_s\}, 1 \leq i_1 < \ldots < i_p < s \) and \( t_1 < t_2 < \cdots < t_p+1 \) such that
\[
f_1, \ldots, f_{i_1} \in \Lambda^t_{k,e} - \Lambda^{t_1-1}_{k,e}, \\
f_{i_1+1}, \ldots, f_{i_2} \in \Lambda^{t_2}_{k,e} - \Lambda^{t_2-1}_{k,e}, \\
\vdots \\
f_{i_{p-1}+1}, \ldots, f_{i_p} \in \Lambda^{t_p}_{k,e} - \Lambda^{t_p-1}_{k,e}, \quad \text{and} \\
f_{i_p+1}, \ldots, f_s \in \Lambda^{t_{p+1}}_{k,e} - \Lambda^{t_{p+1}-1}_{k,e}.
\]

But in such circumstances it is easy to see that
\[
\Lambda^{t_1-1}_{k,e} = \Gamma_{d-1}^{t_1-1} \quad \text{and} \quad \Lambda^{t_2}_{k,e} = \Gamma_d(\Gamma_{d-1}^{t_2-1}) \cup \{f_1, \ldots, f_{i_1}\} = \Gamma_d^{t_2} \cup \{f_1, \ldots, f_{i_1}\}.
\]

Now we can effectively find an index \( q_1 \) such that \( W_{q_1} = \Gamma_{d}^{t_1} \cup \{f_1, \ldots, f_{i_1}\} = \Lambda^{t_1}_{k,e} \) from \( t_1 \) \( f_1, \ldots, f_{i_1} \). Next
\[
\Lambda^{t_2-1}_{k,e} = \Gamma_{d-1}^{t_2-1} (W_{q_1}) \quad \text{and} \quad \Lambda^{t_2}_{k,e} = \Gamma_d(\Gamma_{d-1}^{t_2-1}(W_{q_1})) \cup \{f_{i_1+1}, \ldots, f_{i_2}\}.
\]

Now we can effectively find an index \( q_2 \) such that \( W_{q_2} = \Gamma_{d}^{t_1-1} (W_{q_1}) \cup \{f_{i_1+1}, \ldots, f_{i_2}\} = \Lambda^{t_2}_{k,e} \) from \( q_1, t_2, \) and \( f_{i_1+1}, \ldots, f_{i_2} \). Continuing on this way if we have found an index \( q_r \) such that \( W_{q_{r-1}} = \Lambda^{t_r}_{k,e} \), then
\[
\Lambda^{t_{r-1}}_{k,e} = \Gamma_{d}^{t_{r-1}-1} (W_{q_{r-1}}) \quad \text{and} \quad \Lambda^{t_{r}}_{k,e} = \Gamma_d(\Gamma_{d}^{t_{r-1}-1}(W_{q_{r-1}})) \cup \{f_{r-1+1}, \ldots, f_r\}.
\]

Again, we can effectively find an index \( q_r \) such that
\[
W_{q_r} = \Gamma_d^{t_{r}-t_{r-1}} (W_{q_{r-1}}) \cup \{f_{r-1+1}, \ldots, f_r\} = \Lambda^{t_r}_{k,e}
\]
from \( q_{r-1}, t_r, \) and \( f_{r-1+1}, \ldots, f_r \). Finally, to verify that each stage works properly, we must check that for each \( r \) that \( \{f_{r-1+1}, \ldots, f_r\} \subseteq \Lambda_{k,e}(\Gamma_{d}^{t_{r-1}-1-t_{r-1}}(W_{q_{r-1}})) \) or that for
Inductive Definitions

each $m_j \in \{i_{r-1} + 1, \ldots, i_r\}$. Thus we have established the upper bounds for parts (b) and (c).

For completeness for part (b), fix $\Lambda_{k,e} = \Lambda_{k,e}^3$ and $M_{F,k,e}$ is not closed under $\Lambda_{k,e}$. Now if $F_{k,e} = \Lambda_{k,e}^3$, then clearly $M_{F,k,e}$ is closed under $\Gamma_d$ so that $M_{F,k,e}$ is not closed under $\Lambda_{k,e}$ if and only if

$$\exists m_i \notin F \exists a \in W_e \subseteq M_{F,k,e}$$

which is a $\Sigma_0^3$ predicate. Thus the predicate $|\Lambda_{k,e}| > \omega$ is a conjunction of $\Sigma_0^3$ and $\Pi_0^3$ predicates. Thus we have established the upper bounds for parts (b) and (c).

For the completeness of parts (a), (b), and (c), we will use the $\Sigma_0^3$ complete set $Cof = \{e : W_e$ is cofinite$\}$. Let $P = \{p_0 < p_1 < \cdots \}$ denote the set of primes.

For completeness for part (a), fix $k$ and let $W_{f_i} = \{2^n p_m : n \geq 0$ and $m \geq i\}$ for $i \geq 0$. Then define a 1-1 computable function $g$ so that $(k, g(e)) = (k, (d, \langle m_1, e_1, \ldots, m_k, e_k \rangle))$ where $m_i = i - 1$ and $W_{e_i} = \{f_0, f_1, \ldots\}$, for $i = 1, \ldots, k$, and $\Gamma_d$ is defined so that for all $A \subseteq N$,

1. for all $m \geq k$, $p_m \in \Gamma_d(A) \iff m \in W_e$ and
2. for all $n \geq 1$ and $m \geq k$, $2^n p_m \in \Gamma_d(A) \iff 2^{n-1} p_m \in A$.

It is then easy to see that $\Gamma_1_d = \{p_m : m \in W_e \& m \geq k\}$, $\Gamma_\omega_d = \{2^n p_m : m \in W_e \& m \geq k \& n \geq 0\}$, and there is no finite $t$ such that $W_{f_i} \subseteq \Gamma_\omega_d$ for some $i$. Thus if $W_e$ is cofinite, then there will be an $i$ such $W_{f_i} \subseteq \Gamma_\omega_d$ and, hence, $\{0, \ldots, k-1\} \subseteq \Gamma_\omega_d = \Gamma_\omega_d$. However, if $W_e$ is not cofinite, then there will be no $i$ such that $W_{f_i} \subseteq \Gamma_\omega_d$. Hence $\Gamma_\omega_d = cl(\Lambda_{k,e}(e))$ and $\{0, \ldots, k-1\} \cap cl(\Lambda_{k,e}(e)) = \emptyset$. Thus

$$g(e) \in \{e : (k, e) = (k, (d, \langle m_1, e_1, \ldots, m_k, e_k \rangle)) \& \{m_1, \ldots, m_k\} \cap cl(\Lambda_{k,e} = \emptyset)\}$$

if and only if $W_e$ is not cofinite. It follows that $\{e : (k, e) = (k, (d, \langle m_1, e_1, \ldots, m_k, e_k \rangle)) \& \{m_1, \ldots, m_k\} \cap cl(\Lambda_{k,e} = \emptyset)\}$ is $\Pi_0^3$ complete.

For the completeness for part (b), fix $k$ and for $i = 1, \ldots, k$ let $m_i = i - 1$ and let $W_{e_i} = \{b_0, b_1, b_2, \ldots\}$ where for each $n$, $W_{e_n} = \mathbb{N} - \{0, \ldots, n\}$. Then define the 1:1 computable function $f$ by

$$f(a) = (k, e) = (k, (d, \langle m_1, e_1, \ldots, m_k, e_k \rangle))$$
where $\Gamma_0$ is defined as follows:

For all $A \subseteq \mathbb{N}$,

1. $k \in \Gamma_d(A)$,
2. for all $x \geq 1$, $x + k \in \Gamma_d(A) \iff x \in W_A \lor (\forall y < x) y + k \in A$.

Now if $W_A$ is cofinite, then it is easy to see that $\Lambda_{k,e}^0$ is cofinite and hence $\{0, \ldots, k - 1\} \subseteq \Lambda_{k,e}^2$. It then easily follows that $\Lambda_{k,e}^2 = \mathbb{N}$ and hence $|A_{k,e}| \leq \omega$. However if $W_A$ is not cofinite, then it is easy to see that there is no $t \geq 0$ such that $\Lambda_{k,e}^t$ is cofinite. However it will be the case that $\Lambda_{k,e}^t \ni \{x : k \leq x\}$ and hence, $\Lambda_{k,e}^t = \mathbb{N}$. Thus $a \in \text{Cof} \iff f(a) \in \{e : |A_{k,e}| \leq \omega \text{ and } \{m_1, \ldots, m_k\} \subseteq \Lambda_{k,e}^2\}$ so that $\{e : |A_{k,e}| \leq \omega \text{ and } \{m_1, \ldots, m_k\} \subseteq \Lambda_{k,e}^2\}$ is $\Sigma_3^0$ complete.

For the completeness of part (c), fix $k \geq 2$. Then we need only show that there is 1:1 computable function $h$ such that $h(a,b) \in \{e : |A_{k,e}| \leq \omega\}$ if and only if $W_A$ is cofinite and $W_B$ is not cofinite. Let $P = \{p_0 < p_1 < \ldots\}$ be the set of prime numbers. For each $i$, let $W_{c_i} = \{2^n p_i : n \geq 1\}$. Then let $h$ be the computable function such that $(k, h(a,b)) = (k, (d, \langle m_1, c_1, \ldots, m_k, e_k \rangle))$ where $m_i = 2(i - 1) + 1$ for $i = 1, \ldots, k$, $W_{c_i} = \{b_0, b_1, b_2, \ldots\}$, where $W_{b_i} = \{x + 1 : x \in \mathbb{N}\} - \{1, 3, \ldots, 2i + 1\}$, $W_{c_j} = \{c_0, c_1, c_2, \ldots\}$ for $j = 2, \ldots, k$, where $W_{c_j} = \{2^{n_m} p_m : n \geq 0 \text{ and } m \geq i\}$ and $\Gamma_d$ is defined so that for all $A \subseteq \mathbb{N}$,

1. $2k + 1 \in \Gamma_d(A)$,
2. for all $x \geq 1$, $2(x + k) + 1 \in \Gamma_d(A) \iff x \in W_A \lor (\forall y < x) (2y + k) + 1 \in A$,
3. for all $m \geq 0$, $2p_m \in \Gamma_d(A) \iff m \in W_B$ and
4. for all $m \geq 0$ and $n \geq 2$, $2^{n_m} p_m \in \Gamma_d(A) \iff 2^{n_m - 1} p_m \in A$.

We can use the same analysis that we used in part (a) to conclude $\{2p_m : m \in W_B\} \subseteq \Gamma_d^0$, $\{2^{n_m} p_m : m \in W_B \text{ and } n \geq 1\} \subseteq \Gamma_d^2$, and there is no finite t such that $W_{c_i} \subseteq \Gamma_d^t$ for some i. Moreover, it is the case that $\{3, \ldots, 2k - 1\} \subseteq \Gamma_d^{t - 1} - \Gamma_d^e$ if $W_A$ is cofinite and otherwise $\{3, \ldots, 2k - 1\} \cap \Gamma_d^e = \emptyset$. We can now use our analysis in part (b) to conclude that if $W_A$ is cofinite, then $1 \in \Lambda_{k,e}^0$ and hence $\{1\} \cup \{2 + 1 : s \geq k\} \subseteq \Lambda_{k,e}^0$. However if $W_A$ is not cofinite, then there is no stage $t$ such that $1 \in \Lambda_{k,e}^t$ and hence, $\{2 + 1 : s \geq k\} \subseteq \Lambda_{k,e}^0$ and $1 \in \Lambda_{k,e}^0$. It follows that $|A_{k,h(a,b)}| \leq \omega$ if and only if $W_A$ is cofinite and $W_B$ is not cofinite. Hence for $k \geq 2$, $\{e : |A_{k,e}| \leq \omega\}$ is $D_3^0$ complete.

Next we need to define the family of difference sets of $\Sigma_3^0$ sets. For two $\Sigma_3^0$ sets $A$ and $B$, the difference $A - B$ is the intersection of $\Sigma_3^0$ set and a $\Pi_3^0$ set and is said to be a $2-\Sigma_3^0$ set. For $n > 0$, we say that a set $C$ is $2n-\Sigma_3^0$ if and only if $A$ is the union of $n$ $2-\Sigma_3^0$ sets and is $2n + 1-\Sigma_3$ if and only if $A$ is the union of a $\Sigma_3^0$ set with a $2n-\Sigma_3^0$ set. We say that $A$ is $n-\Pi_3^0$ set if the complement of $A$ is $n-\Sigma_3^0$ set.

We can then prove the following.

**Theorem 4.4.** Fix any computable set $R_t$. Then for each $k$, $\{e : \text{lfp}(A_{k,e}) \cap R_t \text{ is computable}\}$ is a $(2^{k+1} - 1)-\Sigma_3^0$ set.

**Proof.** Fix a set $F \subseteq \{1, \ldots, k\}$. Let $M_{F,k,e} = \Gamma_d^0(\{m_i : i \in F\})$ for each index $(k, e) = (k, (d, \langle m_1, c_1, \ldots, m_k, e_k \rangle))$. We are interested in analyzing the predicate that

$$P(F, k, e) : M_{F,k,e} = \text{lfp}(A_{k,e}) \& R_t \cap M_{F,k,e} \text{ is computable.}$$

(8)
It follows from Lemma 4.1 that \( \text{lfp}(\Lambda_{k,e}) = M_{F,k,e} \) if and only if

1. \( \{m_i : (\exists a \in W_e)(W_a \subseteq M_{F,k,e})\} \subseteq \{m_i : i \in F\} \) and
2. for all \( G \subseteq F \), \( \{m_i : (\exists a \in W_e)(W_a \subseteq M_{G,k,e})\} \nsubseteq \{m_i : i \in G\} \).

The predicate \( \{m_i : (\exists a \in W_e)(W_a \subseteq M_{G,k,e})\} \nsubseteq \{m_i : i \in G\} \) is \( \Sigma^0_3 \) if it follows from Lemma 4.1 that \( (\exists a)(a \in W_e \land W_a \subseteq M_{G,k,e}) \) is \( \Sigma^0_3 \). It follows that the predicate \( \{m_i : (\exists a \in W_e)(W_a \subseteq M_{F,k,e})\} \subseteq \{m_i : i \in F\} \) is \( \Pi^0_2 \) if \( F \neq \{1, \ldots, k\} \). Finally, the predicate \( "M_{F,k,e} \cap R_t" \) is computable in \( \Sigma^0_3 \). Thus if \( F \neq \{1, \ldots, k\} \), the predicate \( P(F, k, e) \) is the conjunction of a \( \Sigma^0_3 \) and \( \Pi^0_2 \) predicate and hence is \( 2\Sigma^0_3 \) predicate. If \( F = \{1, \ldots, k\} \), then, we may omit the \( \Pi^0_2 \) predicate so that \( P(F, k, e) \) is a \( \Sigma^0_3 \) predicate.

It follows that the predicate that \( \{e : \text{lfp}(\Gamma_{k,e}) \cap R_t \text{ is computable}\} \) is a disjunction of \( 2^k - 1 \) \( 2\Sigma^0_3 \) sets and one \( \Sigma^0_3 \) set and hence a \( 2^{k+1} - 1 \) set.

It is important to note that the set of all \( \langle k, e \rangle \) such that \( \text{lfp}(\Lambda_{k,e}) \) itself is computable is just \( \Sigma^0_3 \). In fact, if the set \( R_t \) in Theorem 4.4 is finite or cofinite, then \( \{e : \text{lfp}(\Lambda_{k,e}) \cap R_t \text{ is computable}\} \subseteq \Sigma^0_3 \). That is, for each finite \( F \subseteq \{1, \ldots, k\} \) and each computable set \( R \), the question of whether \( R = M_{F,k,e} \) is a \( \Pi^0_2 \) question since \( M_{F,k,e} \) is uniformly c.e., the predicate \( W_a \subseteq M_{G,k,e} \) is \( \Sigma^0_3 \). It follows that the predicate \( \{m_i : (\exists a \in W_e)(W_a \subseteq M_{F,k,e})\} \subseteq \{m_i : i \in F\} \) is \( \Pi^0_2 \) if \( F \neq \{1, \ldots, k\} \). Finally, the predicate \( "M_{F,k,e} \cap R_t" \) is computable in \( \Sigma^0_3 \). Thus if \( F \neq \{1, \ldots, k\} \), the predicate \( P(F, k, e) \) is the conjunction of a \( \Sigma^0_3 \) and \( \Pi^0_2 \) predicate and hence is \( 2\Sigma^0_3 \) predicate. If \( F = \{1, \ldots, k\} \), then, we may omit the \( \Pi^0_2 \) predicate so that \( P(F, k, e) \) is a \( \Sigma^0_3 \) predicate.

Thus we have the following.

**Theorem 4.5.** \( \{\langle k, e \rangle : \text{lfp}(\Lambda_{k,e}) \text{ is computable}\} \subseteq \Sigma^0_3 \)-complete.

Finally, we give a completeness result for Theorem 4.4 in the case where \( k = 1 \).

**Theorem 4.6.** Let \( R_t \) be a fixed infinite coinfinite computable set. Then \( \{e : \text{lfp}(\Lambda_{1,e}) \cap R_t \text{ is computable}\} \subseteq \Sigma^0_3 \)-complete.

**Proof.** The upper bound on the complexity is given by the proof of Theorem 4.4. For the other direction, fix \( R_t = \{2n : n \in \mathbb{N}\} \) without loss of generality. Let \( C = \{e : \text{lfp}(\Lambda_{1,e}) \cap R_t \text{ is computable}\} \). Note that it is proved in (Soare 1987) that \( \text{Rec} = \{e : W_e \text{ is computable}\} \) and \( \text{Cof} = \{e : W_e \text{ is cofinite}\} \) are \( \Sigma^0_3 \) complete.

For the completeness, first we claim that

\[
D = \{\langle a, b, c \rangle : (W_a \text{ is not cofinite} \land W_b \text{ is computable}) \lor W_c \text{ is computable}\}
\]

is \( 3\Sigma^0_3 \)-complete. That is, let \( S = (B - A) \cup C \), where \( A, B, C \) are \( \Sigma^0_3 \). Then there are functions \( f, g, h \) such that \( a \in A \iff f(a) \in \text{Cof}, b \in B \iff g(b) \in \text{Rec}, \) and...
and \(c \in C \iff h(c) \in \text{Rec}\). Thus \(s = (a, b, c) \in S\) iff \([(f(a) \notin \text{Cof})\) and \(g(b) \in \text{Rec}\) or \(h(c) \in \text{Rec}\)] iff \(\phi(s) = (f(a), g(b), h(c)) \in D\). Thus it suffices to reduce \(D\) to \(C\). That is, we will define a 1-weakly finitary \(\Sigma^0_1\) monotone operator \(\Lambda_{f(a,b,c)}\) such that \(\text{lfp}(\Lambda_{f(a,b,c)}) \cap R_t\) is computable if and only if \(\langle a, b, c \rangle \in D\). Since \(\text{Rec}\) and \(\text{Cof}\) are \(\Sigma^0_3\) complete, it follows that there exists a computable function \(g\) such that \(W_c\) is computable or \(W_a\) is cofinite if and only if \(W_{g(a,c)}\) is cofinite. Let \(h\) be a computable function such that for each \(n, W_{h(n)} = \{8i + 3 : i > n\}\). The 1-weakly finitary inductive operator \(\Lambda = \Lambda_{f(a,b,c)}\) is defined by the following clauses.

1. \(0 \in \Lambda(A)\) if \(W_{h(n)} \subseteq A\) for some \(n\).
2. \(8\langle i, s \rangle + 1 \in \Lambda(A)\) if \(i \in W_{g(a,c), s}\) or \(8\langle i, s + 1 \rangle + 1 \in A\).
3. \(8\langle i, s \rangle + 3 \in \Lambda(A)\) if \(8\langle i, 0 \rangle + 1 \in A\).
4. \(8\langle i, s \rangle + 5 \in \Lambda(A)\) if \(i \in W_{b,s}\) or \(8\langle i, s + 1 \rangle + 5 \in A\).
5. \(8\langle i, 0 \rangle + 2 \in \Lambda(A)\) if \(8\langle i, 0 \rangle + 5 \in A\).
6. \(8\langle i, s \rangle + 7 \in \Lambda(A)\) if \(0 \in A\) and either \(i \in W_{c,s}\) or \(8\langle i, s + 1 \rangle + 7 \in A\).
7. \(8\langle i, 0 \rangle + 4 \in \Lambda(A)\) if \(8\langle i, 0 \rangle + 7 \in A\).
8. \(8\langle i, 0 \rangle + 2 \in \Lambda(A)\) if \(0 \in A\).

It is easy to see that clauses (2)-(8) define a computable monotone inductive operator \(\Lambda\) so that \(\Lambda\) is a 1-weakly finitary \(\Sigma^0_1\) operator where \(S_{\text{A}} = \{0\}\).

Clauses of type (2) and (3) ensure that \(\text{lfp}(\Lambda)\) must include \(\{8i + 3 : i \in W_{g(a,c)}\}\) and clauses of type (4) and (5) ensure that \(\text{lfp}(\Lambda)\) must include \(\{8i + 2 : i \in W_b\}\).

Let \(M = \text{lfp}(\Lambda)\). If \(W_{g(a,c)}\) is cofinite, then one of the clauses of type (1) will apply and then the clauses of type (6), (7), and (8) will ensure that \(M \cap R_t\) equals \(\{0\} \cup \{8i + 2 : i < \omega\} \cup \{8i + 4 : i \in W_c\}\) and, hence, \(M \cap R_t\) will be computable if and only if \(W_c\) is computable. If \(W_{g(a,c)}\) is not cofinite, then \(M \cap R_t\) will consist of \(\{8i + 2 : i \in W_b\}\) and, hence, \(M \cap R_t\) will be computable if and only if \(W_b\) is computable.

If \(\langle a, b, c \rangle \in D\), then there are two cases. First suppose that \(W_c\) is computable. Then \(W_{g(a,c)}\) is cofinite so that \(M \cap R_t\) is computable as desired. Next suppose that \(W_c\) is not computable. Then we must have \(W_a\) not cofinite and \(W_b\) computable. In this case, \(W_{g(a,c)}\) is not cofinite and \(M \cap R_t\) is again computable.

If \(\langle a, b, c \rangle \notin D\), then \(W_c\) is not computable and either \(W_a\) is cofinite or \(W_b\) is not computable. Again there are two cases. First suppose that \(W_a\) is cofinite. Then \(W_{g(a,c)}\) is cofinite, so that \(M \cap R_t\) is not computable, as desired. If \(W_a\) is not cofinite, then \(W_{g(a,c)}\) is not cofinite and \(W_b\) is not computable. Thus again \(M \cap R_t\) is not computable.

We conjecture that a similar completeness result will hold for \(k\)-weakly \(\Sigma^0_1\) operators. Finally, we remark that \(k\)-weakly computable monotone operators may be defined and corresponding versions of Theorems 4.4, 4.5 and 4.6 can be shown.

References

Inductive Definitions


