The Gamma question

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23 June 2016

Coarse computability

Given a set $A \subseteq \mathbb{N}$. How close is A to being computable ?

A recent paradigm : A is coarsely computable. This means there is a computable set R such that the asymptotic density of

$$\{n \colon A(n) = R(n)\}$$

equals 1.

Reference : *Downey, Jockusch, and Schupp, Asymptotic density and computably enumerable sets, Journal of Mathematical Logic, 13, No. 2 (2013)*

The γ-value of a set $A \subseteq \mathbb{N}$

A computable set R tries to approximate a complicated set A :

A : 100100100100 000101001001 010101111010 101010100111
R : 000010110111 010101000101 010001011010 101010100111

≈1/2 correct

≈2/3 correct

≈3/4 correct

≈4/5 correct

Take sup of the asymptotic correctness over all computable R's :

$$\gamma(A) = \sup_{R \text{ computable}} \underline{\rho}\{n\colon\ A(n) = R(n)\}$$

$$\text{where } \underline{\rho}(Z) = \liminf_n \frac{|Z \cap [0,n)|}{n}.$$

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Some examples of values $\gamma(A)$

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$$\text{where } \underline{\rho}(Z) = \liminf_n \frac{|Z \cap [0,n)|}{n}.$$

Some possible values

$$A \text{ computable} \Rightarrow \gamma(A) = 1$$

$$A \text{ random} \Rightarrow \gamma(A) = 1/2.$$

Γ-value of a Turing degree

Andrews, Cai, Diamondstone, Jockusch and Lempp (2013) looked at Turing degrees, rather than sets. They defined

$$\Gamma(A) = \inf\{\gamma(B) : \ B \text{ has the same Turing degree as } A\}$$

A smaller Γ value means that A is further away from computable.

Example

An oracle A is called computably dominated if every function that A computes is below a computable function. *They show :*

- If A is random and computably dominated, then $\Gamma(A) = 1/2$.
- If A is not computably dominated then $\Gamma(A) = 0$.

$\Gamma(A) > 1/2$ implies $\Gamma(A) = 1$

Fact (Hirschfeldt, Jockusch, McNicholl and Schupp)

If $\Gamma(A) > 1/2$ then A is computable (so that $\Gamma(A) = 1$).

The idea is to obtain B of the same Turing degree as A by "padding" :

- "Stretch" the value $A(n)$ over the whole interval $I_n = [(n-1)!, n!)$.
- Since $\gamma(B) > 1/2$ there is a computable R agreeing with B on more than half of the bits in almost every interval I_n.
- So for almost all n, the bit $A(n)$ equals the majority of values $R(k)$ where $k \in I_n$.

The Γ-question

Question (Γ-question, Andrews et al., 2013)

Is there a set $A \subseteq \mathbb{N}$ such that $0 < \Gamma(A) < 1/2$?

$\bullet$	? ? ? ? ? ? ? ? ? ?	$\bullet$	$\times$ $\times$ $\times$ $\times$ $\times$ $\times$ $\times$	$\bullet$
$\Gamma = 0$		$\Gamma = 1/2$		$\Gamma = 1$

Theorem

Let $A \in 2^{\mathbb{N}}$. If $\Gamma(A) < 1/2$ then $\Gamma(A) = 0$.

The proof uses the field of error-correcting codes.

Examples of $\Gamma(A) = 0$: infinitely often equal

We know that $A \subseteq \mathbb{N}$ not computably dominated implies $\Gamma(A) = 0$.

- We say $g : \mathbb{N} \to \mathbb{N}$ is infinitely often equal (i.o.e.) if $\exists^\infty n\ f(n) = g(n)$ for each computable function $f : \mathbb{N} \to \mathbb{N}$.
- We say that $A \subseteq \mathbb{N}$ is i.o.e. if A computes function g that is i.o.e.

Surprising fact : A is i.o.e $\Leftrightarrow$ A not computably dominated.

$\Rightarrow$ Suppose A computes a function g that equals infinitely often to every computable function. Then no computable function bounds g.

$\Leftarrow$ *Idea.* Suppose A computes a function g that is dominated by no computable function. Then g is infinitely often above the halting time of any computable total function.

New Examples of $\Gamma(A) = 0$: weaken infinitely often equal

We know A not computably dominated implies $\Gamma(A) = 0$.

Recall

We say that A is infinitely often equal (i.o.e.) if A computes a function g such that $\exists^\infty n\ f(n) = g(n)$ for each computable function $f : \mathbb{N} \to \mathbb{N}$.

We can weaken this :

Let $H\colon \mathbb{N} \to \mathbb{N}$ be computable. We say that A is H-infinitely often equal if A computes a function g such that $\exists^\infty n\ f(n) = g(n)$ for each computable function f bounded by H.

This appears to get harder for A the faster H grows.

A i.o.e. implies $\Gamma(A) = 0$

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Theorem (Monin, Nies)

Let A be $2^{(\alpha^n)}$-i.o.e. for some $\alpha > 1$. Then $\Gamma(A) = 0$.

New example of $\Gamma(A) = 0$

Recall : A is H-infinitely often equal if A computes a function g such that $\exists^\infty n\ f(n) = g(n)$ for each computable function f bounded by H.

Theorem

Let A be $2^{(\alpha^n)}$-i.o.e. for some computable $\alpha > 1$. Then $\Gamma(A) = 0$.

Proof sketch. First step : Let f be $2^{(\alpha^n)}$-i.o.e. Then for any $k \in \mathbb{N}$, f computes a function g that is $2^{(k^n)}$-i.o.e.

f(0) f(1) f(2) f(3) f(4) f(5) . . . i.o.e. every comp. funct. $\leqslant 2^{(\alpha^n)}$

$\rightarrow$ $f(0)f(2)f(4)\ldots$ i.o.e. every comp. funct. $\leqslant n \mapsto 2^{(\alpha^{2n})}$

or $f(1)f(3)f(5)\ldots$ i.o.e. every comp. funct. $\leqslant n \mapsto 2^{(\alpha^{2n+1})}$

Iterating this $\rightarrow$ $f \geqslant_T g$ which i.o.e. every comp. funct. $\leqslant 2^{(k^n)}$

Proof sketch. Second step : g is $2^{(k^n)}$-i.o.e. implies $g \geqslant_T Z$ with $\Gamma(Z) \leqslant 1/k$.

$$
\begin{array}{rcccccc}
 & g(0) & g(1) & \dots & & g(n) & \dots \\
 & = & = & \dots & & = & \dots \\
Z: & \underbrace{\sigma_0}_{|\sigma_0|=k^0} & \underbrace{\sigma_1}_{|\sigma_1|=k^1} & \dots & & \underbrace{\sigma_n}_{|\sigma_n|=k^n} & \dots \\
\text{Computable } R: & \tau_0 & \tau_1 & \dots & & \tau_n & \dots \\
 & & & \downarrow \text{(bit flip)} & & & \\
\overline{R}: & \overline{\tau_0} & \overline{\tau_1} & \dots & & \overline{\tau_n} & \dots \\
 & = & = & = & & = & \\
 & j(0) & j(1) & \dots & & j(n) & \dots
\end{array}
$$

j equals g infinitely often. Then for infinitely many n, $\tau_n(i) \neq \sigma_n(i)$ everywhere. We have

$$|\tau_n| \geqslant (k-1) \sum_{i<n} |\tau_i|$$

Then the $\liminf$ of fraction of places where R agrees with Z is bounded by $1/k$.

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Nothing between 0 and 1/2

Theorem

Let $X \in \mathbb{N}$. Suppose that for every $k \in \mathbb{N}$ and every X-computable sequence $\{\tau_n\}_{n\in\mathbb{N}}$ with $|\tau_n| = 2^{n/k}$,
there is a computable sequence $\{\sigma_n\}_{n\in\mathbb{N}}$ with $|\sigma_n| = |\tau_n|$ such that for almost every n, σ_n agrees with τ_n on a fraction of at least α bits.
Then $\Gamma(X) \geqslant \alpha$.

Idea : The length of the n-th string equals $2^{1/k} - 1$ times the sum of the length of the previous strings. For c as large as we want, let k be large enough so that $2^{1/k} - 1 < 1/c$.

For $Y \leqslant_T X$, we split Y in strings $\{\tau_n\}_{n\in\mathbb{N}}$ of length $2^{n/k}$. The computable sequence $\{\sigma_n\}_{n\in\mathbb{N}}$ given above implies $\gamma(Y) \geqslant \frac{\alpha}{1+1/c}$.

If this is true for every c we have $\gamma(Y) \geqslant \alpha$. If this is true for every $Y \leqslant_T X$ we have $\Gamma(X) \geqslant \alpha$.

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If this is true for every c we have $\gamma(Y) \geqslant \alpha$. If this is true for every $Y \leqslant_T X$ we have $\Gamma(X) \geqslant \alpha$.

Nothing between 0 and $1/2$

Suppose $\Gamma(X) < 1/2 - \varepsilon$.
Then there is $k \in \mathbb{N}$ and an X-computable sequence $\{\tau_n\}_{n\in\mathbb{N}}$ with $|\tau_n| = 2^{n/k}$, such that :

For every computable sequence $\{\sigma_n\}_{n\in\mathbb{N}}$ with $|\sigma_n| = |\tau_n|$, there are infinitely many n such that σ_n agrees with τ_n on a fraction of **at most** $1/2 - \varepsilon$ *bits.*

By taking the bitwise complement of every such computable sequence $\{\sigma_n\}_{n\in\mathbb{N}}$ we get :

For every computable sequence $\{\sigma_n\}_{n\in\mathbb{N}}$ with $|\sigma_n| = |\tau_n|$, there are infinitely many n such that σ_n agrees with τ_n on a fraction of **at least** $1/2 + \varepsilon$ *bits.*

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The error-correcting codes

We want to transmit a message of length m on a noisy chanel. We use an injection $\Phi : 2^m \to 2^n$ for $n > m$ in such a way that the strings in the range of Φ are pairwise as far as possible.

If δ is the smallest relative Hamming distance between two strings in the range of Φ, we can correct up to a fraction of $\delta/2$ errors.

We cannot in general correct more than a ratio of $1/4$ of errors. To go beyond we need to use List decoding :

Theorem (List decoding theorem)

Let $\varepsilon > 0$ and $n \in \mathbb{N}$. For $L \in \mathbb{N}$ sufficiently large and $\beta > 0$ sufficiently small, there exists a set C of $2^{\beta n}$ many strings of length n such that :
For any string σ of length n, there are at most L elements τ of C such that σ agrees with τ on a fraction of bits of at least $1/2 + \varepsilon$.

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Then there is $k \in \mathbb{N}$ and an X-computable sequence $\{\tau_n\}_{n\in\mathbb{N}}$ with $|\tau_n| = 2^{n/k}$, such that :
For every computable sequence $\{\sigma_n\}_{n\in\mathbb{N}}$ with $|\sigma_n| = |\tau_n|$, there are infinitely many n such that σ_n agrees with τ_n on a fraction of at least $1/2 + \varepsilon$ bits.

For any n we compute a sequence C_n of $2^{(\beta 2^{n/k})}$ many strings of length $2^{n/k}$ such that any string σ of length $2^{n/k}$ agrees with at most L elements of C_n on a fraction of at least $1/2 + \varepsilon$ bits.

From $\{\tau_n\}_{n\in\mathbb{N}}$, we compute the sequence $\{D_n\}_{n\in\mathbb{N}}$ of all the strings of length $\beta 2^{n/k}$ whose code in C_n agrees with τ_n on more than $1/2 + \varepsilon$ bits. We have $|D_n| \leqslant L$ for every n.

Claim : *For every computable function g bounded by $2^{(\beta 2^{n/k})}$, there are infinitely many n such that $g(n) \in D_n$ (seen as a binary string).*

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From this we compute :
An X-computable sequence $\{D_n\}_{n\in\mathbb{N}}$ where D_n contains at most L strings of length $L2^n$ and such that :
For every computable function g bounded by $2^{(L2^n)}$, there are infinitely many n such that $g(n) \in D_n$ (seen as a binary string).

We see the i-th element σ_i of D_n as an L-uplet $\langle \sigma_i^1, \dots, \sigma_i^L \rangle$. Let h_i be the function which to n gives σ_i^i where σ_i is the i-th string of D_n.

At least one h_i must be $2^{(2^n)}$-i.o.e., which concludes the proof.

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